Contents

2 Programming Language Syntax

2.3.5 Syntax Errors

EXAMPLE 2.43

Syntax error in C (reprise)

The main text illustrated the problem of syntax error recovery with a simple example in C:

```
A = B : C + D;
```

The compiler will detect a syntax error immediately after the `B`, but it cannot give up at that point: it needs to keep looking for errors in the remainder of the program. To permit this, we must modify the input program, the state of the parser, or both, in a way that allows parsing to continue, hopefully without announcing a significant number of spurious cascading errors and without missing a significant number of real errors. The techniques discussed below allow the compiler to search for further syntax errors. In Chapter 4 we will consider additional techniques that allow it to search for additional static semantic errors as well. ■

Panic Mode

Perhaps the simplest form of syntax error recovery is a technique known as *panic mode*. It defines a small set of "safe symbols" that delimit clean points in the input. When an error occurs, a panic mode recovery algorithm deletes input tokens until it finds a safe symbol, then backs the parser out to a context in which that symbol might appear. In the earlier example, a recursive descent parser with panic mode recovery might delete input tokens until it finds the semicolon, return from all subroutines called from within stmt, and restart the body of stmt itself.

Unfortunately, panic mode tends to be a bit drastic. By limiting itself to a static set of "safe" symbols at which to resume parsing, it admits the possibility of deleting a significant amount of input while looking for such a symbol. Worse, if some of the deleted tokens are "starter" symbols that begin large-scale constructs in the language (e.g., `begin`, `procedure`, `while`), we shall almost surely see spurious cascading errors when we reach the end of the construct.

EXAMPLE 2.44

The problem with panic mode

Consider the following fragment of code in an Algol-family language:

```
IF a b THEN x;
ELSE y;
END;
```

When it discovers the error at `b` in the first line, a panic-mode recovery algorithm is likely to skip forward to the semicolon, thereby missing the `THEN`. When the parser finds the `ELSE` on line 2 it will produce a spurious error message. When it finds the `END` on line 3 it will think it has reached the end of the enclosing structure (e.g., the whole subroutine), and will probably generate additional cascading errors on subsequent lines. Panic mode tends to work acceptably only in relatively "unstructured" languages, such as Basic and (early) Fortran, which don't have many "starter" symbols. ■

Phrase-Level Recovery

We can improve the quality of recovery by employing different sets of "safe" symbols in different contexts. Parsers that incorporate this improvement are said to implement *phrase-level recovery*. When it discovers an error in an expression, for example, a phrase-level recovery algorithm can delete input tokens until it reaches something that is likely to follow an expression. This more local recovery is better than always backing out to the end of the current statement, because it gives us the opportunity to examine the parts of the statement that follow the erroneous expression.

EXAMPLE 2.45
Phrase-level recovery in recursive descent

Niklaus Wirth, the inventor of Pascal, published an elegant implementation of phrase-level recovery for recursive descent parsers in 1976 [Wir76, Sec. 5.9]. The simplest version of his algorithm depends on the FIRST and FOLLOW sets defined at the end of Section 2.3.1. If the parsing routine for nonterminal *foo* discovers an error at the beginning of its code, it deletes incoming tokens until it finds a member of FIRST(*foo*), in which case it proceeds, or a member of FOLLOW(*foo*), in which case it returns:

```
procedure foo()
    if not (input_token ∈ FIRST(foo) or EPS(foo))
        report_error()                  –– print message for the user
        repeat
            delete_token()
        until input_token ∈ (FIRST(foo) ∪ FOLLOW(foo) ∪ {$$})
    case input_token of
        . . . : . . .
        . . . : . . .                   –– valid starting tokens
        . . . : . . .
        otherwise return                –– error or foo ⟶ ϵ
```

Note that the report_error routine does *not* terminate the parse; it simply prints a message and returns. To complete the algorithm, the match routine must be altered so that it, too, will return after announcing an error, effectively inserting the expected token when something else appears:

```
procedure match(expected)
    if input_token = expected
        consume_input_token()
    else
        report_error()
```

Finally, to simplify the code, the common prefix of the various nonterminal subroutines can be moved into an error-checking subroutine:

```
procedure check_for_error(symbol)
    if not (input_token $\in$ FIRST(symbol) or EPS(symbol))
        report_error()
        repeat
            delete_token()
        until input_token $\in$ (FIRST(symbol) $\cup$ FOLLOW(symbol) $\cup$ {$$})
```
■

Context-Specific Look-Ahead

Though simple, the recovery algorithm just described has an unfortunate tendency, when *foo* $\longrightarrow \epsilon$, to predict one or more epsilon productions when it should really announce an error right away. This weakness is known as the *immediate error detection* problem. It stems from the fact that FOLLOW(*foo*) is context-independent: it contains all tokens that may follow *foo* somewhere in some valid program, but not necessarily in the current context in the current program. This is basically the same observation that underlies the distinction between SLR and LALR parsers ("The Characteristic Finite-State Machine and LR Parsing Variants," Section 2.3.4).

EXAMPLE 2.46
Cascading syntax errors

As an example, consider the following incorrect code in our calculator language:

```
Y := (A * X X*X) + (B * X*X) + (C * X) + D
```

To a human being, it is pretty clear that the programmer forgot a `*` in the x^3 term of a polynomial. The recovery algorithm isn't so smart. In a recursive descent parser it will see an identifier (`X`) coming up on the input when it is inside the following routines:

```
program
stmt_list
stmt
expr
term
factor
expr
term
factor_tail
factor_tail
```

Since an `id` can follow a *factor_tail* in some programs (e.g., `A := B  C := D`), the innermost parsing routine will predict *factor_tail* $\longrightarrow \epsilon$, and simply return. At that point both the outer factor_tail and the inner term will be at the end of their code, and they, too, will return. Next, the inner expr will call term_tail, which will also predict an epsilon production, since an `id` can follow a term_tail in certain programs. This will leave the inner expr at the end of its code, allowing it to return. Only then will we discover an error, when factor calls match, expecting to see a right parenthesis. Afterward there will be a host of cascading errors, as the input is transformed into

```
Y := (A * X)
X := X
B := X*X
C := X
```

■

EXAMPLE 2.47
Reducing cascading errors with context-specific look-ahead

To avoid inappropriate epsilon predictions, Wirth introduced the notion of context-specific FOLLOW sets, passed into each nonterminal subroutine as an explicit parameter. In our example, we would pass `id` as part of the FOLLOW set for the initial, outer expr, which is called as part of the production *stmt* $\longrightarrow$ `id :=` *expr*, but *not* into the second, inner expr, which is called as part of the production *factor* $\longrightarrow$ `(` *expr* `)`. The nested calls to term and factor_tail will end up being called with a FOLLOW set whose only member is a right parenthesis. When the inner call to factor_tail discovers that `id` is not in FIRST(*factor_tail*), it will delete tokens up to the right parenthesis before returning. The net result is a single error message, and a transformation of the input into

```
Y := (A * X*X) + (B * X*X) + (C * X) + D
```

That's still not the "right" interpretation, but it's a lot better than it was. ■

The final version of Wirth's phrase-level recovery employs one additional heuristic: to avoid cascading errors it refrains from deleting members of a statically defined set of "starter" symbols (e.g., `begin`, `procedure`, `(`, etc.). These are the symbols that tend to require matching tokens later in the program. If we see a starter symbol while deleting input, we give up on the attempt to delete the rest of the erroneous construct. We simply return, even though we know that the starter symbol will not be acceptable to the calling routine. With context-specific FOLLOW sets and starter symbols, phrase-level recovery looks like this:

EXAMPLE 2.48
Recursive descent with full phrase-level recovery

```
procedure check_for_error(symbol, follow_set)
    if not (input_token ∈ FIRST(symbol) or (EPS(symbol) and input_token ∈ follow_set))
        report_error()
        repeat
            delete_token()
        until input_token ∈ FIRST(symbol) ∪ follow_set ∪ starter_set ∪ {$$}
```

```
procedure expr(follow_set)
    check_for_error(expr, follow_set)
    case input_token of
        . . . : . . .
        . . . : . . .                valid starting tokens
        . . . : . . .
        otherwise return
```

■

Exception-Based Recovery in Recursive Descent

An attractive alternative to Wirth's technique relies on the exception-handling mechanisms available in many modern languages (we will discuss these mechanisms in detail in Section 9.4). Rather than implement recovery for every nonterminal in the language (a somewhat tedious task), the exception-based approach identifies a small set of contexts to which we back out in the event of an error. In many languages, we could obtain simple, but probably serviceable error recovery by backing out to the nearest statement or declaration. In the limit, if we choose a single place to "back out to," we have an implementation of panic-mode recovery.

EXAMPLE 2.49
Exceptions in a recursive descent parser

The basic idea is to attach an exception handler (a special syntactic construct) to the blocks of code in which we want to implement recovery:

```
procedure statement()
    try
        . . .                   –– code to parse a statement
    except when syntax_error
        loop
            if next_token ∈ FIRST(statement)
                statement()         –– try again
                return
            elsif next_token ∈ FOLLOW(statement)
                return
            else get_next_token()
```

Code for declaration would be similar. For better-quality repair, we might add handlers around the bodies of expression, aggregate, or other complex constructs. To guarantee that we can always recover from an error, we must ensure that all parts of the grammar lie inside at least one handler.

When we detect an error (possibly nested many procedure calls deep), we *raise* a syntax error exception ("`raise`" is a built-in command in languages with exceptions). The language implementation then unwinds the stack to the most recent context in which we have an exception handler, which it executes in place of the remainder of the block to which the handler is attached. For phrase-level (or panic mode) recovery, the handler can delete input tokens until it sees one with which it can recommence parsing. ■

As noted in Section 2.3.1, the ANTLR parser generator takes a CFG as input and builds a human-readable recursive descent parser. Compiler writers have the option of generating Java, C#, or C++, all of which have exception-handling mechanisms. When an ANTLR-generated parser encounters a syntax error, it throws a MismatchedTokenException or NoViableAltException. By default ANTLR includes a handler for these exceptions in every nonterminal subroutine. The handler prints an error message, deletes tokens until it finds something in the FOLLOW set of the nonterminal, and then returns. The compiler writer can define alternative handlers if desired on a production-by-production basis.

Error Productions

As a general rule, it is desirable for an error recovery technique to be as language-independent as possible. Even in a recursive descent parser, which is handwritten for a particular language, it is nice to be able to encapsulate error recovery in the check_for_error and match subroutines. Sometimes, however, one can obtain much better repairs by being highly language specific.

EXAMPLE 2.50
Error production for "; else"

Most languages have a few unintuitive rules that programmers tend to violate in predictable ways. In Pascal, for example, semicolons are used to separate statements, but many programmers think of them as *terminating* statements instead. Most of the time the difference is unimportant, since a statement is allowed to be empty. In the following, for example,

```
begin
    x := (-b + sqrt(b*b -4*a*c)) / (2*a);
    writeln(x);
end;
```

the compiler parses the begin... end block as a sequence of three statements, the third of which is empty. In the following, however,

```
if d <> 0 then
    a := n/d;
else
    a := n;
end;
```

the compiler must complain, since the then part of an if... then ... else construct must consist of a single statement in Pascal. A Pascal semicolon is never allowed immediately before an else, but programmers put them there all the time. Rather than try to tune a general recovery or repair algorithm to deal correctly with this problem, most Pascal compiler writers modify the grammar: they include an extra production that allows the semicolon, but causes the semantic analyzer to print a warning message, telling the user that the semicolon shouldn't be there. Similar error productions are used in C compilers to cope with "anachronisms" that have crept into the language as it evolved. Syntax that was valid only in early versions of C is still accepted by the parser, but evokes a warning message. ■

Error Recovery in Table-Driven LL Parsers

Given the similarity to recursive descent parsing, it is straightforward to implement phrase-level recovery in a table-driven top-down parser. Whenever we encounter an error entry in the parse table, we simply delete input tokens until we find a member of a statically defined set of starter symbols (including `$$`), or a member of the FIRST or FOLLOW set of the nonterminal at the top of the parse stack.[1] If we find a member of the FIRST set, we continue the main loop of the driver. If we find a member of the FOLLOW set or the starter set, we pop the nonterminal off the parse stack first. If we encounter an error in `match`, rather than in the parse table, we simply pop the token off the parse stack.

But we can do better than this! Since we have the entire parse stack easily accessible (it was hidden in the control flow and procedure calling sequence of recursive descent), we can enumerate all possible combinations of insertions and deletions that would allow us to continue parsing. Given appropriate metrics, we can then evaluate the alternatives to pick the one that is in some sense "best."

Because perfect error recovery (actually error *repair*) would require that we read the programmer's mind, any practical technique to evaluate alternative "corrections" must rely on heuristics. For the sake of simplicity, most compilers limit themselves to heuristics that (1) require no semantic information, (2) do not require that we "back up" the parser or the input stream (i.e., to some state prior to the one in which the error was detected), and (3) do not change the spelling of tokens or the boundaries between them. A particularly elegant algorithm that conforms to these limits was published by Fischer, Milton, and Quiring in 1980 [FMQ80]. As originally described, the algorithm was limited to languages in which programs could always be corrected by inserting appropriate tokens into the input stream, without ever requiring deletions. It is relatively easy, however, to extend the algorithm to encompass deletions and substitutions. We consider the insert-only algorithm first; the version with deletions employs it as a subroutine. We do not consider substitutions here.[2]

The FMQ error-repair algorithm requires the compiler writer to assign an insertion cost $C(\texttt{t})$ and a deletion cost $D(\texttt{t})$ to every token `t`. (Since we cannot change where the input ends, we have $C(\texttt{\$\$}) = D(\texttt{\$\$}) = \infty$.) In any given error situation, the algorithm chooses the least cost combination of insertions and

1 This description uses global FOLLOW sets. If we want to use context-specific look-aheads instead, we can peek farther down in the stack. A token is an acceptable context-specific look-ahead if it is in the FIRST set of the second symbol A from the top in the stack or, if it would cause us to predict $A \longrightarrow \epsilon$, the FIRST set of the third symbol B from the top or, if it would cause us to predict $B \longrightarrow \epsilon$, the FIRST set of the fourth symbol from the top, and so on.

2 A substitution can always be effected as a deletion/insertion pair, but we might want ideally to give it special consideration. For example, we probably want to be cautious about deleting a left square bracket or inserting a left parenthesis, since both of these symbols must be matched by something later in the input, at which point we are likely to see cascading errors. But substituting a left parenthesis for a left square bracket is in some sense more plausible, especially given the differences in array subscript syntax in different programming languages.

deletions that allows the parser to consume one more token of real input. The state of the parser is never changed; only the input is modified (rather than pop a stack symbol, the repair algorithm pushes its yield onto the input stream).

As in phrase-level recovery in a recursive descent parser, the FMQ algorithm needs to address the immediate error detection problem. There are several ways we could do this. One would be to use a "full LL" parser, which keeps track of local FOLLOW sets. Another would be to inspect the stack when predicting an epsilon production, to see if what lies underneath will allow us to accept the incoming token. The first option significantly increases the size and complexity of the parser. The second option leads to a nonlinear-time parsing algorithm. Fortunately, there is a third option. We can save all changes to the stack (and calls to the semantic analyzer's action routines) in a temporary buffer until the match routine accepts another real token of input. If we discover an error before we accept a real token, we undo the stack changes and throw away the buffered calls to action routines. Then we can pretend we recognized the error when a full LL parser would have.

We now consider the task of repairing with only insertions. We begin by extending the notion of insertion costs to strings in the obvious way: if $w = a_1 a_2 \ldots a_n$, we have $C(w) = \sum_{i=1}^{n} C(a_i)$. Using the cost function C, we then build a pair of tables S and E. The S table is one-dimensional, and is indexed by grammar symbol. For any symbol X, $S(X)$ is a least-cost string of terminals derivable from X. That is,

$$S(X) = w \iff X \Longrightarrow^* w \text{ and } \forall x \text{ such that } X \Longrightarrow^* x,\ C(w) \le C(x)$$

Clearly $S(\texttt{a}) = \texttt{a}\ \forall$ tokens a.

The E table is two-dimensional, and is indexed by symbol/token pairs. For any symbol X and token a, $E(X, \texttt{a})$ is the lowest-cost prefix of a in X; that is, the lowest cost token string w such that $X \Longrightarrow^* w\texttt{a}x$. If X cannot yield a string containing a, then $E(X, \texttt{a})$ is defined to be a special symbol ?? whose insertion cost is ∞. If $X = \texttt{a}$, or if $X \Longrightarrow^* \texttt{a}x$, then $E(X, \texttt{a}) = \epsilon$, where $C(\epsilon) = 0$.

EXAMPLE 2.51
Insertion-only repair in FMQ

To find a least-cost insertion that will repair a given error, we execute the function find_insertion, shown in Figure C-2.31. The function begins by considering the least-cost insertion that will allow it to derive the input token from the symbol at the top of the stack (there may be none). It then considers the possibility of "deleting" that top-of-stack symbol (by inserting its least-cost yield into the input stream) and deriving the input token from the second symbol on the stack. It continues in this fashion, considering ways to derive the input token from ever deeper symbols on the stack, until the cost of inserting the yields of the symbols above exceeds the cost of the cheapest repair found so far. If it reaches the bottom of the stack without finding a finite-cost repair, then the error cannot be repaired by insertions alone. ■

EXAMPLE 2.52
FMQ with deletions

To produce better-quality repairs, and to handle languages that cannot be repaired with insertions only, we need to consider deletions. As we did with the insert cost vector C, we extend the deletion cost vector D to strings of tokens in the obvious way. We then embed calls to find_insertion in a second loop, shown

```
function find_insertion(a : token) : string
    –– assume that the parse stack consists of symbols $X_n, \ldots X_2, X_1$,
    –– with $X_n$ at top-of-stack
    ins := ??
    prefix := $\epsilon$
    for $i$ in $n \,..\, 1$
        if $C$(prefix) $\geq C$(ins)
            –– no better insertion is possible
            return ins
        if $C$(prefix . $E(X_i$, a)) < $C$(ins)
            –– better insertion found
            ins := prefix . $E(X_i$, a)
        prefix := prefix . $S(X_i)$
    return ins
```

Figure 2.31 Outline of a function to find a least-cost insertion that will allow the parser to accept the input token *a*. The dot character (.) is used here for string concatenation.

```
function find_repair() : ⟨string, int⟩
    –– assume that the parse stack consists of symbols $X_n, \ldots X_2, X_1$,
    –– with $X_n$ at top-of-stack,
    –– and that the input stream consists of tokens $a_1, a_2, a_3, \ldots$
    i := 0        –– number of tokens we're considering deleting
    best_ins := ??
    best_del := 0
    loop
        cur_ins := find_insertion($a_{i+1}$)
        if $C$(cur_ins) + $D(a_1 \ldots a_i)$ < $C$(best_ins) + $D(a_1 \ldots a_{\text{best\_del}})$
            –– better repair found
            best_ins := cur_ins
            best_del := i
        i +:= 1
        if $D(a_1 \ldots a_i) > C$(best_ins) + $D(a_1 \ldots a_{\text{best\_del}})$
            –– no better repair is possible
            return ⟨best_ins, best_del⟩
```

Figure 2.32 Outline of a function to find a least-cost combination of insertions and deletions that will allow the parser to accept one more token of input.

in Figure C-2.32. This loop repeatedly considers deleting more and more tokens, each time calling find_insertion on the remaining input, until the cost of deleting additional tokens exceeds the cost of the cheapest repair found so far. The search can never fail; it is always possible to find a combination of insertions and deletions that will allow the end-of-file token to be accepted. Since the algorithm may need to consider (and then reject) the option of deleting an arbitrary number of tokens, the scanner must be prepared to peek an arbitrary distance ahead in the input stream and then back up again. ■

The FMQ algorithm has several desirable properties. It is simple and efficient (given that the grammar is bounded in size, we can prove that the time to choose a repair is bounded by a constant). It can repair an arbitrary input string. Its decisions are locally optimal, in the sense that no cheaper repair can allow the parser to make forward progress. It is table-driven and therefore fully automatic. Finally, it can be tuned to prefer "likely" repairs by modifying the insertion and deletion costs of tokens. Some obvious heuristics include:

- Deletion should usually be more expensive than insertion.
- Common operators (e.g., multiplication) should have lower cost than uncommon operators (e.g., modulo division) in the same place in the grammar.
- Starter symbols (e.g., `begin`, `if`, `(`) should have higher cost than their corresponding final symbols (`end`, `fi`, `)`).
- "Noise" symbols (comma, semicolon, `do`) should have very low cost.

Error Recovery in Bottom-Up Parsers

Locally least-cost repair is possible in bottom-up parsers, but it isn't as easy as it is in top-down parsers. The advantage of a top-down parser is that the content of the parse stack unambiguously identifies the context of an error, and specifies the constructs expected in the future. The stack of a bottom-up parser, by contrast, describes a set of possible contexts, and says nothing explicit about the future.

In practice, most bottom-up parsers tend to rely on panic-mode or phrase-level recovery. The intuition is that when an error occurs, the top few states on the parse stack represent the shifted prefix of an erroneous construct. Recovery consists of popping these states off the stack, deleting the remainder of the construct from the incoming token stream, and then restarting the parser, possibly after shifting a fictitious nonterminal to represent the erroneous construct.

Unix's `yacc/bison` provides a typical example of bottom-up phrase-level recovery. In addition to the usual tokens of the language, `yacc/bison` allows the compiler writer to include a special token, `error`, anywhere in the right-hand sides of grammar productions. When the parser built from the grammar detects a syntax error, it

1. Calls the function `yyerror`, which the compiler writer must provide. Normally, `yyerror` simply prints a message (e.g., "parse error"), which `yacc/bison` passes as an argument
2. Pops states off the parse stack until it finds a state in which it can shift the `error` token (if there is no such state, the parser terminates)
3. Inserts and then shifts the `error` token
4. Deletes tokens from the input stream until it finds a valid look-ahead for the new (post `error`) context
5. Temporarily disables reporting of further errors
6. Resumes parsing

If there are any semantic action routines associated with the production containing the `error` token, these are executed in the normal fashion. They can do such things as print additional error messages, modify the symbol table, patch up semantic processing, prompt the user for additional input in an interactive tool (`yacc/bison` can be used to build things other than batch-mode compilers), or disable code generation. The rationale for disabling further syntax errors is to make sure that we have really found an acceptable context in which to resume parsing before risking cascading errors. `Yacc/bison` automatically reenables the reporting of errors after successfully shifting three real tokens of input. A semantic action routine can reenable error messages sooner if desired by calling the built-in routine `yyerrorok`.

EXAMPLE 2.53
Panic mode in `yacc/bison`

For our example calculator language, we can imagine building a `yacc/bison` parser using the bottom-up grammar of Figure 2.25. For panic-mode recovery, we might want to back out to the nearest statement:

stmt $\longrightarrow$ `error`
`{printf("parsing resumed at end of current statement\n");}`

The semantic routine written in curly braces would be executed when the parser recognizes *stmt* $\longrightarrow$ `error`.[3] Parsing would resume at the next token that can follow a statement—in our calculator language, at the next `id`, `read`, `write`, or `$$`. ■

EXAMPLE 2.54
Panic mode with statement terminators

A weakness of the calculator language, from the point of view of error recovery, is that the current, erroneous statement may well contain additional `id`s. If we resume parsing at one of these, we are likely to see another error right away. We could avoid the error by disabling error messages until several real tokens have been shifted. In a language in which every statement ends with a semicolon, we could have more safely written

stmt $\longrightarrow$ `error ;`
`{printf("parsing resumed at end of current statement\n");}` ■

EXAMPLE 2.55
Phrase-level recovery in `yacc/bison`

In both of these examples we have placed the `error` symbol at the beginning of a right-hand side, but there is no rule that says it must be so. We might decide, for example, that we will abandon the current statement whenever we see an error, unless the error happens inside a parenthesized expression, in which case we will attempt to resume parsing after the closing parenthesis. We could then add the following production:

factor $\longrightarrow$ `( error )`
`{printf("parsing resumed at end of parenthesized expression\n");}`

3 The syntax shown here is not the same as that accepted by `yacc/bison`, but is used for the sake of consistency with earlier material.

In the CFSM of Figure 2.26, it would then be possible in State 8 to shift `error`, delete some tokens, shift `)`, recognize *factor*, and continue parsing the surrounding expression. Of course, if the erroneous expression contains nested parentheses, the parser may not skip all of it, and a cascading error may still occur. ■

Because `yacc/bison` creates LALR parsers, it automatically employs context-specific look-ahead, and does not usually suffer from the immediate error detection problem. (A full LR parser would do slightly better.) In an SLR parser, a good error recovery algorithm needs to employ the same trick we used in the top-down case. Specifically, we buffer all stack changes and calls to semantic action routines until the shift routine accepts a real token of input. If we discover an error before we accept a real token, we undo the stack changes and throw away the buffered calls to semantic routines. Then we can pretend we recognized the error when a full LR parser would have.

✓ CHECK YOUR UNDERSTANDING

45. Why is syntax error recovery important?
46. What are *cascading errors*?
47. What is *panic mode*? What is its principal weakness?
48. What is the advantage of *phrase-level recovery* over panic mode?
49. What is the *immediate error detection problem*, and how can it be addressed?
50. Describe two situations in which context-specific FOLLOW sets may be useful.
51. Outline Wirth's mechanism for error recovery in recursive descent parsers. Compare this mechanism to exception-based recovery.
52. What are *error productions*? Why might a parser that incorporates a high-quality, general-purpose error recovery algorithm still benefit from using such productions?
53. Outline the FMQ algorithm. In what sense is the algorithm optimal?
54. Why is error recovery more difficult in bottom-up parsers than it is in top-down parsers?
55. Describe the error recovery mechanism employed by `yacc/bison`.

2 Programming Language Syntax

2.4 Theoretical Foundations

As noted in the main text, scanners and parsers are based on the finite automata and pushdown automata that form the bottom two levels of the Chomsky language hierarchy. At each level of the hierarchy, machines can be either *deterministic* or *nondeterministic*. A deterministic automaton always performs the same operation in a given situation. A nondeterministic automaton can perform any of a *set* of operations. A nondeterministic machine is said to accept a string if there exists a choice of operation in each situation that will eventually lead to the machine saying "yes." As it turns out, nondeterministic and deterministic finite automata are equally powerful: every DFA is, by definition, a degenerate NFA, and the construction in Example 2.14 demonstrates that for any NFA we can create a DFA that accepts the same language. The same is not true of push-down automata: there are context-free languages that are accepted by an NPDA but not by any DPDA. Fortunately, DPDAs suffice in practice to accept the syntax of real programming languages. Practical scanners and parsers are always deterministic.

2.4.1 Finite Automata

Precisely defined, a deterministic finite automaton (DFA) M consists of (1) a finite set Q of *states*, (2) a finite alphabet Σ of input symbols, (3) a distinguished *initial* state $q_1 \in Q$, (4) a set of distinguished *final* states $F \subseteq Q$, and (5) a *transition function* $\delta : Q \times \Sigma \rightarrow Q$ that chooses a new state for M based on the current state and the current input symbol. M begins in state q_1. One by one it consumes its input symbols, using δ to move from state to state. When the final symbol has been consumed, M is interpreted as saying "yes" if it is in a state in F; otherwise it is interpreted as saying "no." We can extend δ in the obvious way to take strings, rather than symbols, as inputs, allowing us to say that M accepts string x if $\delta(q_1, x) \in F$. We can then define $L(M)$, the language accepted by M, to be the set

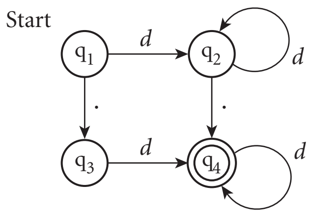


Figure 2.33 **Minimal DFA for the language consisting of all strings of decimal digits containing a single decimal point.** Adapted from Figure 2.10 in the main text. The symbol d here is short for "`0, 1, 2, 3, 4, 5, 6, 7, 8, 9`".

$\{x \mid \delta(q_1, x) \in F\}$. In a nondeterministic finite automaton (NFA), the transition function, δ, is multivalued: the automaton can move to any of a *set* of possible states from a given state on a given input. In addition, it may move from one state to another "spontaneously"; such transitions are said to take input symbol ϵ.

EXAMPLE 2.56
Formal DFA for
d*(. d | d .) d*

We can illustrate these definitions with an example. Consider the circles-and-arrows automaton of Figure C-2.33 (adapted from Figure 2.10 in the main text). This is the minimal DFA accepting strings of decimal digits containing a single decimal point. $\Sigma = \{\texttt{0}, \texttt{1}, \texttt{2}, \texttt{3}, \texttt{4}, \texttt{5}, \texttt{6}, \texttt{7}, \texttt{8}, \texttt{9}, \texttt{.}\}$ is the machine's input alphabet. $Q = \{q_1, q_2, q_3, q_4\}$ is the set of states; q_1 is the initial state; $F = \{q_4\}$ (a singleton in this case) is the set of final states. The transition function can be represented by a set of triples $\delta = \{(q_1, \texttt{0}, q_2), \ldots, (q_1, \texttt{9}, q_2), (q_1, \texttt{.}, q_3), (q_2, \texttt{0}, q_2), \ldots, (q_2, \texttt{9}, q_2), (q_2, \texttt{.}, q_4), (q_3, \texttt{0}, q_4), \ldots, (q_3, \texttt{9}, q_4), (q_4, \texttt{0}, q_4), \ldots, (q_4, \texttt{9}, q_4)\}$. In each triple $(q_i, \texttt{a}, q_j)$, $\delta(q_i, \texttt{a}) = q_j$. ■

Given the constructions of Examples 2.12 and 2.14, we know that there exists an NFA that accepts the language generated by any given regular expression, and a DFA equivalent to any given NFA. To show that regular expressions and finite automata are of equivalent expressive power, all that remains is to demonstrate that there exists a regular expression that generates the language accepted by any given DFA. We illustrate the required construction below for our decimal strings example (Figure C-2.33). More formal and general treatment of all the regular language constructions can be found in standard automata theory texts [HMU07, Sip13].

From a DFA to a Regular Expression

To construct a regular expression equivalent to a given DFA, we employ a dynamic programming algorithm that builds solutions to successively more complicated subproblems from a table of solutions to simpler subproblems. We begin with a set of simple regular expressions that describe the transition function, δ. For all states i, we define

$$r_{ii}^0 = \texttt{a}_1 \mid \texttt{a}_2 \mid \ldots \mid \texttt{a}_m \mid \epsilon$$

where $\{\texttt{a}_1 \mid \texttt{a}_2 \mid \ldots \mid \texttt{a}_m\} = \{\texttt{a} \mid \delta(q_i, \texttt{a}) = q_i\}$ is the set of characters labeling the "self-loop" from state q_i back to itself. If there is no such self-loop, $r^0_{ij} = \epsilon$. Similarly, for $i \neq j$, we define

$$r^0_{ij} = \texttt{a}_1 \mid \texttt{a}_2 \mid \ldots \mid \texttt{a}_m$$

where $\{\texttt{a}_1 \mid \texttt{a}_2 \mid \ldots \mid \texttt{a}_m\} = \{\texttt{a} \mid \delta(q_i, \texttt{a}) = q_j\}$ is the set of characters labeling the arc from q_i to q_j. If there is no such arc, r^0_{ij} is the empty regular expression. (Note the difference here: we can stay in state q_i by not accepting any input, so ϵ is always one of the alternatives in r^0_{ii}, but not in r^0_{ij} when $i \neq j$.)

Given these r^0 expressions, the dynamic programming algorithm inductively calculates expressions r^k_{ij} with larger superscripts. In each, k names the highest-numbered state through which control may pass on the way from q_i to q_j. We assume that states are numbered starting with q_1, so when $k = 0$ we must transition directly from q_i to q_j, with no intervening states.

EXAMPLE 2.57

Reconstructing a regular expression for the decimal string DFA

In our small example DFA, $r^0_{11} = r^0_{33} = \epsilon$, and $r^0_{22} = r^0_{44} = \texttt{0} \mid \texttt{1} \mid \texttt{2} \mid \texttt{3} \mid \texttt{4} \mid \texttt{5} \mid \texttt{6} \mid \texttt{7} \mid \texttt{8} \mid \texttt{9} \mid \epsilon$, which we will abbreviate $d \mid \epsilon$. Similarly, $r^0_{13} = r^0_{24} = \texttt{.}$, and $r^0_{12} = r^0_{34} = d$. Expressions r^0_{14}, r^0_{21}, r^0_{23}, r^0_{31}, r^0_{32}, r^0_{41}, r^0_{42}, and r^0_{43} are all empty.

For $k > 0$, the r^k_{ij} expressions will generally generate multicharacter strings. At each step of the dynamic programming algorithm, we let

$$r^k_{ij} = r^{k-1}_{ij} \mid r^{k-1}_{ik} r^{k-1}_{kk}{}^{*} r^{k-1}_{kj}$$

That is, to get from q_i to q_j without going through any states numbered higher than k, we can either go from q_i to q_j without going through any state numbered higher than $k-1$ (which we already know how to do), or else we can go from q_i to q_k (without going through any state numbered higher than $k-1$), travel out from q_k and back again an arbitrary number of times (never visiting a state numbered higher than $k-1$ in between), and finally go from q_k to q_j (again without visiting a state numbered higher than $k-1$). If any of the constituent regular expressions is empty, we omit its term of the outermost alternation. At the end, our overall answer is $r^n_{1f_1} \mid r^n_{1f_2} \mid \ldots \mid r^n_{1f_t}$, where $n = |Q|$ is the total number of states and $F = \{q_{f_1}, q_{f_2}, \ldots, q_{f_t}\}$ is the set of final states.

Because $r^0_{11} = \epsilon$ and there are no transitions from States 2, 3, or 4 to State 1, nothing changes in the first inductive step in our example; that is, $\forall i\,[r^1_{ii} = r^0_{ii}]$. The second step is a bit more interesting. Since we are now allowed to go through State 2, we have $r^2_{22} = r^2_{22}\ r^2_{22}{}^{*} r^2_{22} = (\,d \mid \epsilon\,) \mid (\,d \mid \epsilon\,)\,(\,d \mid \epsilon\,)^{*}(\,d \mid \epsilon\,) = d^{*}$. Because r^1_{21}, r^1_{23}, r^1_{32}, and r^1_{42} are empty, however, r^2_{11}, r^2_{33}, and r^2_{44} remain the same as r^1_{11}, r^1_{33}, and r^1_{44}. In a similar vein, we have

$$\begin{aligned} r^2_{12} &= d \mid d\,(\,d \mid \epsilon\,)^{*}(\,d \mid \epsilon\,) \;=\; d^{+} \\ r^2_{14} &= d\,(\,d \mid \epsilon\,)^{*}\,\texttt{.} \;=\; d^{+}\,\texttt{.} \\ r^2_{24} &= \texttt{.} \mid (\,d \mid \epsilon\,)\,(\,d \mid \epsilon\,)^{*}\,\texttt{.} \;=\; d^{*}\,\texttt{.} \end{aligned}$$

Missing transitions and empty expressions from the previous step leave $r^2_{13} = r^1_{13}$ = . and $r^2_{34} = r^1_{34} = d$. Expressions $r^2_{21}, r^2_{23}, r^2_{31}, r^2_{32}, r^2_{41}, r^2_{42}$, and r^2_{43} remain empty.

In the third inductive step, we have

$$\begin{aligned} r^3_{13} &= \texttt{.} \mid \texttt{.}\,\epsilon^\star\epsilon \;=\; \texttt{.} \\ r^3_{14} &= d^+\,\texttt{.} \mid \texttt{.}\,\epsilon^\star d \;=\; d^+\,\texttt{.} \mid \texttt{.}\,d \\ r^3_{34} &= d \mid \epsilon\epsilon^\star d \;=\; d \end{aligned}$$

All other expressions remain unchanged from the previous step.

Finally, we have

$$\begin{aligned} r^4_{14} &= (\,d^+\,\texttt{.} \mid \texttt{.}\,d\,) \mid (\,d^+\,\texttt{.} \mid \texttt{.}\,d\,)\,(\,d \mid \epsilon\,)^\star(\,d \mid \epsilon\,) \\ &= (\,d^+\,\texttt{.} \mid \texttt{.}\,d\,) \mid (\,d^+\,\texttt{.} \mid \texttt{.}\,d\,)\,d^\star \\ &= (\,d^+\,\texttt{.} \mid \texttt{.}\,d\,)\,d^\star \\ &= d^+\,\texttt{.}\,d^\star \mid \texttt{.}\,d^+ \end{aligned}$$

Since F has a single member (q_4), this expression is our final answer. ■

Space Requirements

In Section 2.2.1 we noted without proof that the conversion from an NFA to a DFA may lead to exponential blow-up in the number of states. Certainly this did not happen in our decimal string example: the NFA of Figure 2.8 has 14 states, while the equivalent DFA of Figure 2.9 has only 7, and the minimal DFA of Figures 2.10 and C-2.33 has only 4.

EXAMPLE 2.58
A regular language with a large minimal DFA

Consider, however, the subset of (a | b | c)* in which some letter appears at least three times. The minimal DFA for this language has 28 states. As shown in Figure C-2.34, 27 of these are states in which we have seen i, j, and k as, bs, and cs, respectively. The 28th (and only final) state is reached once we have seen at least three of some specific character.

By contrast, there exists an NFA for this language with only eight states, as shown in Figure C-2.35. It requires that we "guess," at the outset, whether we will see three as, three bs, or three cs. It mirrors the structure of the natural regular expression (a | b | c)* a (a | b | c)* a (a | b | c)* a (a | b | c)* | (a | b | c)* b (a | b | c)* b (a | b | c)* b (a | b | c)* | (a | b | c)* c (a | b | c)* c (a | b | c)* c (a | b | c)*. ■

EXAMPLE 2.59
Exponential DFA blow-up

Of course, the eight-state NFA does not emerge directly from the construction of Figure 2.7; that construction produces a 52-state machine with a certain amount of redundancy, and with many extraneous states and epsilon productions. But consider the similar subset of (0 | 1 | 2 | 3 | 4 | 5 | 6 | 7 | 8 | 9)* in which some digit appears at least ten times. The minimal DFA for this language has 10,000,000,001 states: a non-final state for each combination of zeros through nines with less than ten of each, and a single final state reached once any digit has appeared at least ten times. One possible regular expression for this language is

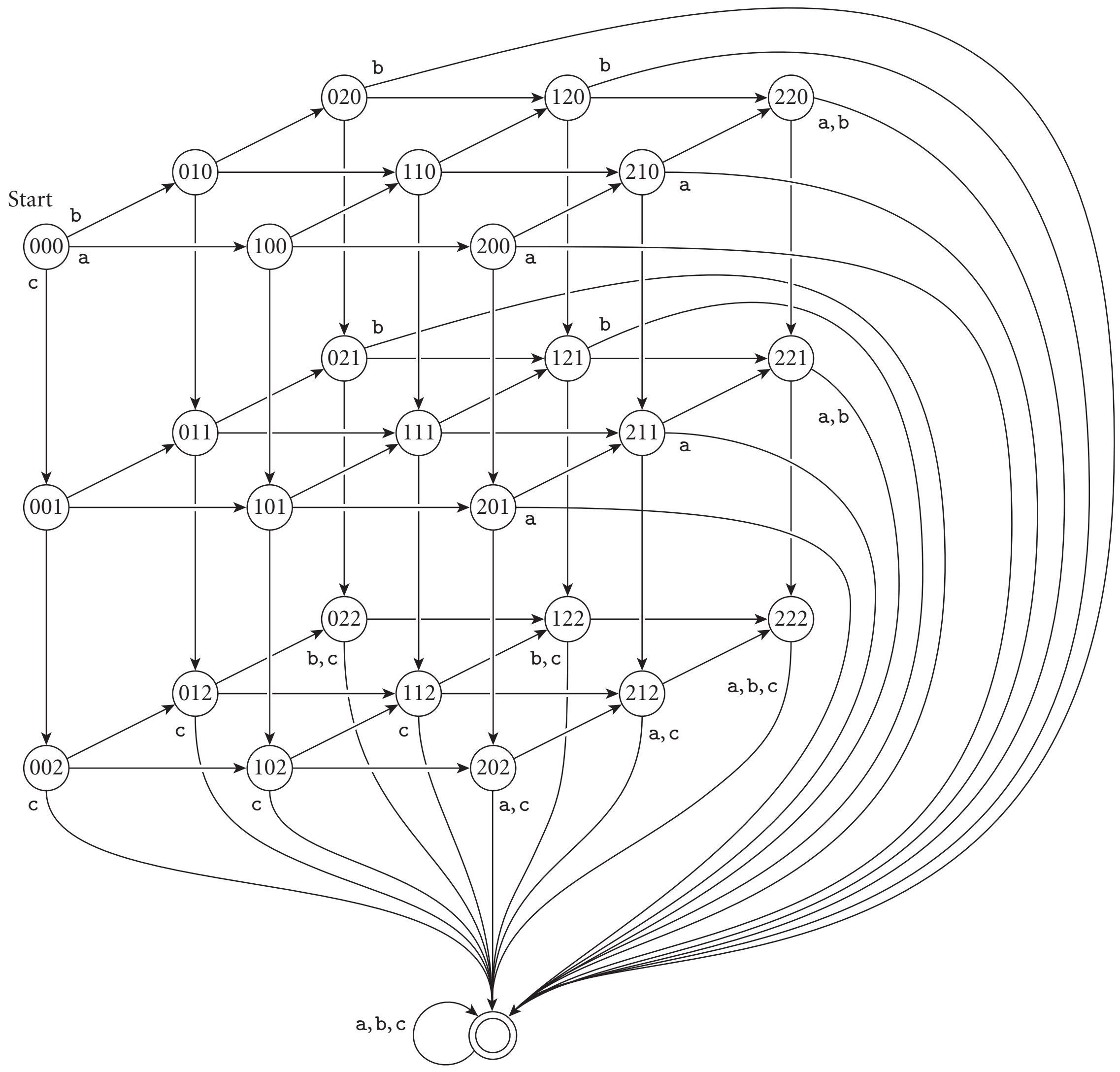


Figure 2.34 **DFA for the language consisting of all strings in `( a | b | c )*` in which some letter appears at least three times.** State name *ijk* indicates that *i* `a`s, *j* `b`s, and *k* `c`s have been seen so far. Within the cubic portion of the figure, most edge labels are elided: `a` transitions move to the right, `b` transitions go back into the page, and `c` transitions move down.

```
 ((0|1|...|9)* 0 (0|1|...|9)* 0 (0|1|...|9)* 0 (0|1|...|9)* 0
   (0|1|...|9)* 0 (0|1|...|9)* 0 (0|1|...|9)* 0 (0|1|...|9)* 0
   (0|1|...|9)* 0 (0|1|...|9)* 0 (0|1|...|9)*)
|((0|1|...|9)* 1 (0|1|...|9)* 1 (0|1|...|9)* 1 (0|1|...|9)* 1
   (0|1|...|9)* 1 (0|1|...|9)* 1 (0|1|...|9)* 1 (0|1|...|9)* 1
   (0|1|...|9)* 1 (0|1|...|9)* 1 (0|1|...|9)*)
|  ...
|((0|1|...|9)* 9 (0|1|...|9)* 9 (0|1|...|9)* 9 (0|1|...|9)* 9
   (0|1|...|9)* 9 (0|1|...|9)* 9 (0|1|...|9)* 9 (0|1|...|9)* 9
   (0|1|...|9)* 9 (0|1|...|9)* 9 (0|1|...|9)*)
```

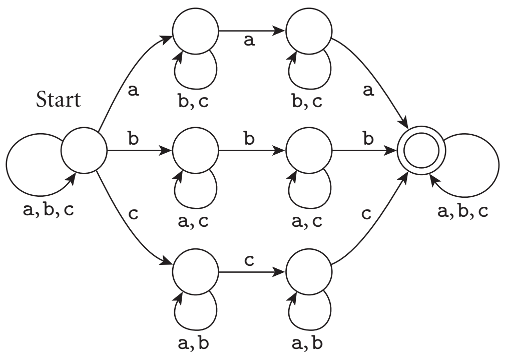


Figure 2.35 NFA corresponding to the DFA of Figure C-2.34.

Our construction would yield a very large NFA for this expression, but clearly many orders of magnitude smaller than ten billion states! ■

2.4.2 Push-Down Automata

A deterministic push-down automaton (DPDA) N consists of (1) Q, (2) Σ, (3) q_1, and (4) F, as in a DFA, plus (6) a finite alphabet Γ of stack symbols, (7) a distinguished initial stack symbol $Z_1 \in \Gamma$, and (5′) a transition function $\delta : Q \times \Gamma \times \{\Sigma \cup \{\epsilon\}\} \rightarrow Q \times \Gamma^*$, where Γ^* is the set of strings of zero or more symbols from Γ. N begins in state q_1, with symbol Z_1 in an otherwise empty stack. It repeatedly examines the current state q and top-of-stack symbol Z. If $\delta(q,\epsilon,Z)$ is defined, N moves to state r and replaces Z with α in the stack, where $(r,\alpha) = \delta(q,\epsilon,Z)$. In this case N does not consume its input symbol. If $\delta(q,\epsilon,Z)$ is undefined, N examines and consumes the current input symbol a. It then moves to state s and replaces Z with β, where $(s,\beta) = \delta(q,\texttt{a},Z)$. N is interpreted as accepting a string of input symbols if and only if it consumes the symbols and ends in a state in F.

As with finite automata, a nondeterministic push-down automaton (NPDA) is distinguished by a multivalued transition function: an NPDA can choose any of a set of new states and stack symbol replacements when faced with a given state, input, and top-of-stack symbol. If $\delta(q,\epsilon,Z)$ is nonempty, N can also choose a new state and stack symbol replacement without inspecting or consuming its current input symbol. While we have seen that nondeterministic and deterministic finite automata are equally powerful, this correspondence does not carry over to push-down automata: there are context-free languages that are accepted by an NPDA but not by any DPDA.

The proof that CFGs and NPDAs are equivalent in expressive power is more complex than the corresponding proof for regular expressions and finite automata. The proof is also of limited practical importance for compiler construction; we do not present it here. While it is possible to create an NPDA for any

CFL, simulating that NPDA may in some cases require exponential time to recognize strings in the language. (The $O(n^3)$ algorithms mentioned in Section 2.3 do not take the form of PDAs.) Practical programming languages can all be expressed with LL or LR grammars, which can be parsed with a (deterministic) PDA in linear time.

An LL(1) PDA is very simple. Because it makes decisions solely on the basis of the current input token and top-of-stack symbol, its state diagram is trivial. All but one of the transitions is a self-loop from the initial state to itself. A final transition moves from the initial state to a second, final state when it sees `$$` on the input and the stack. As we noted in Section 2.3.4, the state diagram for an SLR(1) or LALR(1) parser is substantially more interesting: it's the characteristic finite-state machine (CFSM). Full LR(1) parsers have similar machines, but usually with many more states, due to the need for path-specific look-ahead.

A little study reveals that if we define every state to be accepting, then the CFSM, without its stack, is a DFA that recognizes the grammar's *viable prefixes*. These are all the strings of grammar symbols that can begin a sentential form in the canonical (right-most) derivation of some string in the language, and that do not extend beyond the end of the handle. The algorithms to construct LL(1) and SLR(1) PDAs from suitable grammars were given in Sections 2.3.3 and 2.3.4.

2.4.3 Grammar and Language Classes

EXAMPLE 2.60
0^n1^n is not a regular language

As we noted in Section 2.1.2, a scanner is incapable of recognizing arbitrarily nested constructs. The key to the proof is to realize that we cannot count an arbitrary number of left-bracketing symbols with a finite number of states. Consider, for example, the problem of accepting the language 0^n1^n. Suppose there is a DFA M that accepts this language. Suppose further that M has m states. Now suppose we feed M a string of $m+1$ zeros. By the *pigeonhole principle* (you can't distribute m objects among $p < m$ pigeonholes without putting at least two objects in some pigeonhole), M must enter some state q_i twice while scanning this string. Without loss of generality, let us assume it does so after seeing j zeros and again after seeing k zeros, for $j \neq k$. Since we know that M accepts the string 0^j1^j and the string 0^k1^k, and since it is in precisely the same state after reading 0^j and 0^k, we can deduce that M must also accept the strings 0^j1^k and 0^k1^j. Since these strings are not in the language, we have a contradiction: M cannot exist. ■

Within the family of context-free languages, one can prove similar theorems about the constructs that can and cannot be recognized using various parsing algorithms. Though almost all real parsers get by with a single token of look-ahead, it is possible in principle to use more than one, thereby expanding the set of grammars that can be parsed in linear time. In the top-down case we can redefine FIRST and FOLLOW sets to contain pairs of tokens in a more or less straightforward fashion. If we do this, however, we encounter a more serious version of the immediate error detection problem described in Section C-2.3.5. There we saw that the use of context-independent FOLLOW sets could cause us to overlook a syntax error until after we had needlessly predicted one or more epsilon productions. Context-specific FOLLOW sets solved the problem, but did not change the set of *valid* programs that could be parsed with one token of look-ahead. If we define LL(k) to be the set of all grammars that can be parsed predictively using the top-of-stack symbol and k tokens of look-ahead, then it turns out that for $k > 1$ we must adopt a context-specific notion of FOLLOW sets in order to parse correctly. The algorithm of Section 2.3.3, which is based on context-independent FOLLOW sets, is actually known as SLL (simple LL), rather than true LL. For $k = 1$, the LL(1) and SLL(1) algorithms can parse the same set of grammars. For $k > 1$, LL is strictly more powerful. Among the bottom-up parsers, the relationships among SLR(k), LALR(k), and LR(k) are somewhat more complicated, but extra look-ahead always helps.

EXAMPLE 2.61
Separation of grammar classes

Containment relationships among the classes of grammars accepted by popular linear-time algorithms appear in Figure C-2.36. The LR class (no suffix) contains every grammar G for which there exists a k such that $G \in$ LR(k); LL, SLL, SLR, and LALR are similarly defined. Grammars can be found in every region of the figure. Examples appear in Figure C-2.37. Proofs that they lie in the regions claimed are deferred to Exercise C-2.35. ■

For any context-free grammar G and parsing algorithm P, we say that G is a P grammar (e.g., an LL(1) grammar) if it can be parsed using that algorithm. By extension, for any context-free *language* L, we say that L is a P language if there exists a P grammar for L (this may not be the grammar we were given).

EXAMPLE 2.62
Separation of language classes

Containment relationships among the classes of languages accepted by the popular parsing algorithms appear in Figure C-2.38. Again, languages can be found in every region. Examples appear in Figure C-2.39; proofs are deferred to Exercise C-2.36. ■

Note that every context-free language that can be parsed deterministically has an SLR(1) grammar. Moreover, any language that can be parsed deterministically and in which no valid string can be extended to create another valid string (this is called the *prefix property*) has an LR(0) grammar. If we restrict our attention to languages with an explicit `$$` marker at end-of-file, then they all have the prefix property, and therefore LR(0) grammars.

The relationships among language classes are not as rich as the relationships among grammar classes. Most real programming languages can be parsed by any of the popular parsing algorithms, though the grammars are not always pretty, and special-purpose “hacks” may sometimes be required when a language is almost, but not quite, in a given class. The principal advantage of the more powerful parsing algorithms (e.g., full LR) is that they can parse a wider variety of grammars for a given language. In practice this flexibility makes it easier for the compiler writer to find a grammar that is intuitive and readable, and that facilitates the creation of semantic action routines.

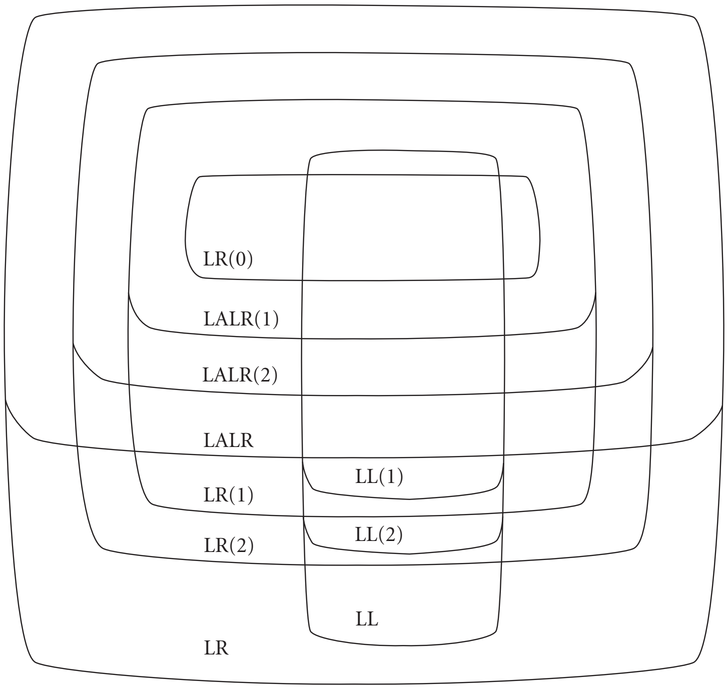


Figure 2.36 **Containment relationships among popular grammar classes.** In addition to the containments shown, SLL(k) is just inside LL(k), for $k \geq 2$, but has the same relationship to everything else, and SLR(k) is just inside LALR(k), for $k \geq 1$, but has the same relationship to everything else.

LL(2) but not SLL:

$$S \longrightarrow \texttt{a}\ A\ \texttt{a} \mid \texttt{b}\ A\ \texttt{b}\ \texttt{a}$$
$$A \longrightarrow \texttt{b} \mid \epsilon$$

SLL(k) but not LL($k-1$):

$$S \longrightarrow \texttt{a}^{k-1}\ \texttt{b} \mid \texttt{a}^{k}$$

LR(0) but not LL:

$$S \longrightarrow A\ \texttt{b}$$
$$A \longrightarrow A\ \texttt{a} \mid \texttt{a}$$

SLL(1) but not LALR:

$$S \longrightarrow A\ \texttt{a} \mid B\ \texttt{b} \mid \texttt{c}\ C$$
$$C \longrightarrow A\ \texttt{b} \mid B\ \texttt{a}$$
$$A \longrightarrow D$$
$$B \longrightarrow D$$
$$D \longrightarrow \epsilon$$

SLL(k) and SLR(k) but not LR($k-1$):

$$S \longrightarrow A\ \texttt{a}^{k-1}\ \texttt{b} \mid B\ \texttt{a}^{k-1}\ \texttt{c}$$
$$A \longrightarrow \epsilon$$
$$B \longrightarrow \epsilon$$

LALR(1) but not SLR:

$$S \longrightarrow \texttt{b}\ A\ \texttt{b} \mid A\ \texttt{c} \mid \texttt{a}\ \texttt{b}$$
$$A \longrightarrow \texttt{a}$$

LR(1) but not LALR:

$$S \longrightarrow \texttt{a}\ C\ \texttt{a} \mid \texttt{b}\ C\ \texttt{b} \mid \texttt{a}\ D\ \texttt{b} \mid \texttt{b}\ D\ \texttt{a}$$
$$C \longrightarrow \texttt{c}$$
$$D \longrightarrow \texttt{c}$$

Unambiguous but not LR:

$$S \longrightarrow \texttt{a}\ S\ \texttt{a} \mid \epsilon$$

Figure 2.37 Examples of grammars in various regions of Figure C-2.36.

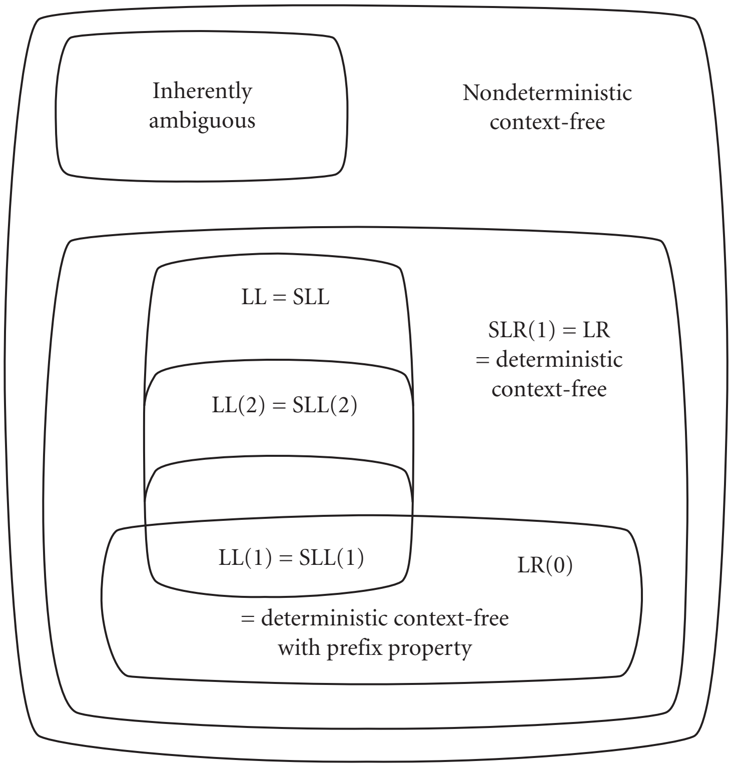

Inherently ambiguous
Nondeterministic context-free
LL = SLL
SLR(1) = LR = deterministic context-free
LL(2) = SLL(2)
LL(1) = SLL(1)
LR(0)
= deterministic context-free with prefix property

Figure 2.38 Containment relationships among popular language classes.

Nondeterministic language:
$\{\texttt{a}^n\texttt{b}^n\texttt{c} : n \geq 1\} \cup \{\texttt{a}^n\texttt{b}^{2n}\texttt{d} : n \geq 1\}$

Inherently ambiguous language:
$\{\texttt{a}^i\texttt{b}^j\texttt{c}^k : i = j \text{ or } j = k;\ i, j, k \geq 1\}$

Language with LL(k) grammar but no LL(k−1) grammar:
$\{\texttt{a}^n(\texttt{b} \mid \texttt{c} \mid \texttt{b}^k\texttt{d})^n : n \geq 1\}$

Language with LR(0) grammar but no LL grammar:
$\{\texttt{a}^n\texttt{b}^n : n \geq 1\} \cup \{\texttt{a}^n\texttt{c}^n : n \geq 1\}$

Figure 2.39 Examples of languages in various regions of Figure C-2.38.

CHECK YOUR UNDERSTANDING

56. What formal machine captures the behavior of a scanner? A parser?
57. State three ways in which a real scanner differs from the formal machine.
58. What are the formal components of a DFA?

59. Outline the algorithm used to construct a regular expression equivalent to a given DFA.

60. What is the inherent "big-O" complexity of parsing with a simulated NPDA? Why is this worse than the $O(n^3)$ time mentioned in Section 2.3?

61. How many states are there in an LL(1) PDA? An SLR(1) PDA? Explain.

62. What are the *viable prefixes* of a CFG?

63. Summarize the proof that a DFA cannot recognize arbitrarily nested constructs.

64. Explain the difference between LL and SLL parsing.

65. Is every LL(1) grammar also LR(1)? Is it LALR(1)?

66. Does every LR language have an SLR(1) grammar?

67. Why are the containment relationships among grammar classes more complex than those among language classes?

2 Programming Language Syntax

2.6 Exercises

2.31 Give an example of an erroneous program fragment in which consideration of semantic information (e.g., types) might help one make a good choice between two plausible "corrections" of the input.

2.32 Give an example of an erroneous program fragment in which the "best" correction would require one to "back up" the parser (i.e., to undo recent predictions/matches or shifts/reductions).

2.33 Extend your solution to exercise 2.21 to implement Wirth's syntax error recovery mechanism

(a) with global FOLLOW sets, as in Example C-2.45.

(b) with local FOLLOW sets, as in Example C-2.47.

(c) with avoidance of "starter symbol" deletion, as in Example C-2.48.

2.34 Extend your solution to exercise 2.21 to implement exception-based syntax error recovery, as in Example C-2.49.

2.35 Prove that the grammars in Figure C-2.37 lie in the regions claimed.

2.36 (Difficult) Prove that the languages in Figure C-2.39 lie in the regions claimed.

2.37 Prove that regular expressions and *left-linear grammars* are equally powerful. A left-linear grammar is a context-free grammar in which every right-hand side contains at most one nonterminal, and then only at the left-most end.

2 Programming Language Syntax

2.7 Explorations

2.46 Experiment with syntax errors in your favorite compiler. Feed the compiler deliberate errors and comment on the quality of the recovery or repair. How often does it do the "right thing"? How often does it generate cascading errors? Speculate as to what sort of recovery or repair algorithm it might be using.

2.47 Spelling mistakes (typos in keywords and identifiers) are a common source of syntax and static semantic errors. Identifying such errors—and guessing what the user meant to type—could result in significantly better error recovery. Discuss how you might go about incorporating spelling correction into some existing error recovery system. (Hint: You might want to consult Morgan's early paper on this subject [Mor70].)

3 Names, Scopes, and Bindings

3.4 Implementing Scope

For both static and dynamic scoping, a language implementation must keep track of the name-to-object bindings in effect at each point in the program. The principal difference is *time*: with static scope the compiler uses a *symbol table* to track bindings at compile time; with dynamic scoping the interpreter or run-time system uses an *association list* or *central reference table* to track bindings at run time.

3.4.1 Symbol Tables

In a language with static scoping, the compiler uses an insert operation to place a name-to-object binding into the symbol table for each newly encountered declaration. When it encounters the use of a name that should already have been declared, the compiler uses a lookup operation to search for an existing binding. It is tempting to try to accommodate the visibility rules of static scoping by performing a remove operation to delete a name from the symbol table at the end of its scope. Unfortunately, several factors make this straightforward approach impractical:

- The ability of inner declarations to hide outer ones in most languages with nested scopes means that the symbol table has to be able to contain an arbitrary number of mappings for a given name. The lookup operation must return the innermost mapping, and outer mappings must become visible again at end of scope.
- Records (structures) and classes have some of the properties of scopes, but do not share their nicely nested structure. When it sees a record declaration, the semantic analyzer must remember the names of the record's fields (recursively, if records are nested). At the end of the declaration, the field names must become invisible. Later, however, whenever a variable of the record type appears in the program text (as in my_rec.field_name), the record fields must

suddenly become visible again for the part of the reference after the dot. In object-oriented languages, member (field and method) names must become visible throughout the methods of the class, even if (as in C++) the code for the methods can appear outside the class declaration.

- As noted in Section 3.3.3, names are sometimes used before they are declared. Algol and C, for example, permit *forward references* to labels. Pascal permits forward references in pointer declarations. Most object-oriented languages permit forward references to class members. Modula-3 permits forward references of all kinds.
- As noted in Section 3.3.3, C, C++, and Ada distinguish between the declaration of an object and its definition. Pascal has a similar mechanism for mutually recursive subroutines. When it sees a declaration, the compiler must remember any nonvisible details so that it can check the eventual definition for consistency. This operation is similar to remembering the field names of records and classes.
- While it may be desirable to forget names at the end of their scope, and even to reclaim the space they occupy in the symbol table, information about them may need to be saved for use by a *symbolic debugger* (Section 16.3.2). A debugger allows the user to manipulate a running program: starting it, stopping it, and reading and writing its data. In order to parse high-level commands, the debugger must have access to the compiler's symbol table, which the compiler typically saves in a hidden portion of the final machine-language program.

EXAMPLE 3.45
The LeBlanc-Cook symbol table

To accommodate these concerns, most compilers never delete anything from the symbol table. Instead, they manage visibility using enter_scope and leave_scope operations. Implementations vary from compiler to compiler; the approach described here is due to LeBlanc and Cook [CL83].

Each scope, as it is encountered, is assigned a serial number. The outermost scope (the one that contains the predefined identifiers) is given number 0. The scope containing programmer-declared global names is given number 1. Additional scopes are given successive numbers as they are encountered. All serial numbers are distinct; they do not represent the level of lexical nesting, except in as much as nested subroutines naturally end up with numbers higher than those of surrounding scopes.

All names, regardless of scope, are entered into a single large hash table, keyed by name. Each entry in the table then contains the symbol name, its category (variable, constant, type, procedure, field name, parameter, etc.), scope number, type (a pointer to another symbol table entry), and additional, category-specific fields.

In addition to the hash table, the symbol table has a *scope stack* that indicates, in order, the scopes that compose the current referencing environment. As the semantic analyzer scans the program, it pushes and pops this stack whenever it enters or leaves a scope, respectively. Entries in the scope stack contain the scope number, an indication of whether the scope is closed, and in some cases further information.

```
procedure lookup(name)
    pervasive := best := null
    apply hash function to name to find appropriate chain
    foreach entry e on chain
        if e.name = name        –– not something else with same hash value
            if e.scope = 0
                pervasive := e
            else
                foreach scope s on scope stack, top first
                    if s.scope = e.scope
                        best := e        –– closer instance
                        exit inner loop
                    elsif best ≠ null and then s.scope = best.scope
                        exit inner loop        –– won't find better
                    if s.closed
                        exit inner loop        –– can't see farther
    if best ≠ null
        while best is an import or export entry
            best := best.real_entry
        return best
    elsif pervasive ≠ null
        return pervasive
    else
        return null        –– name not found
```

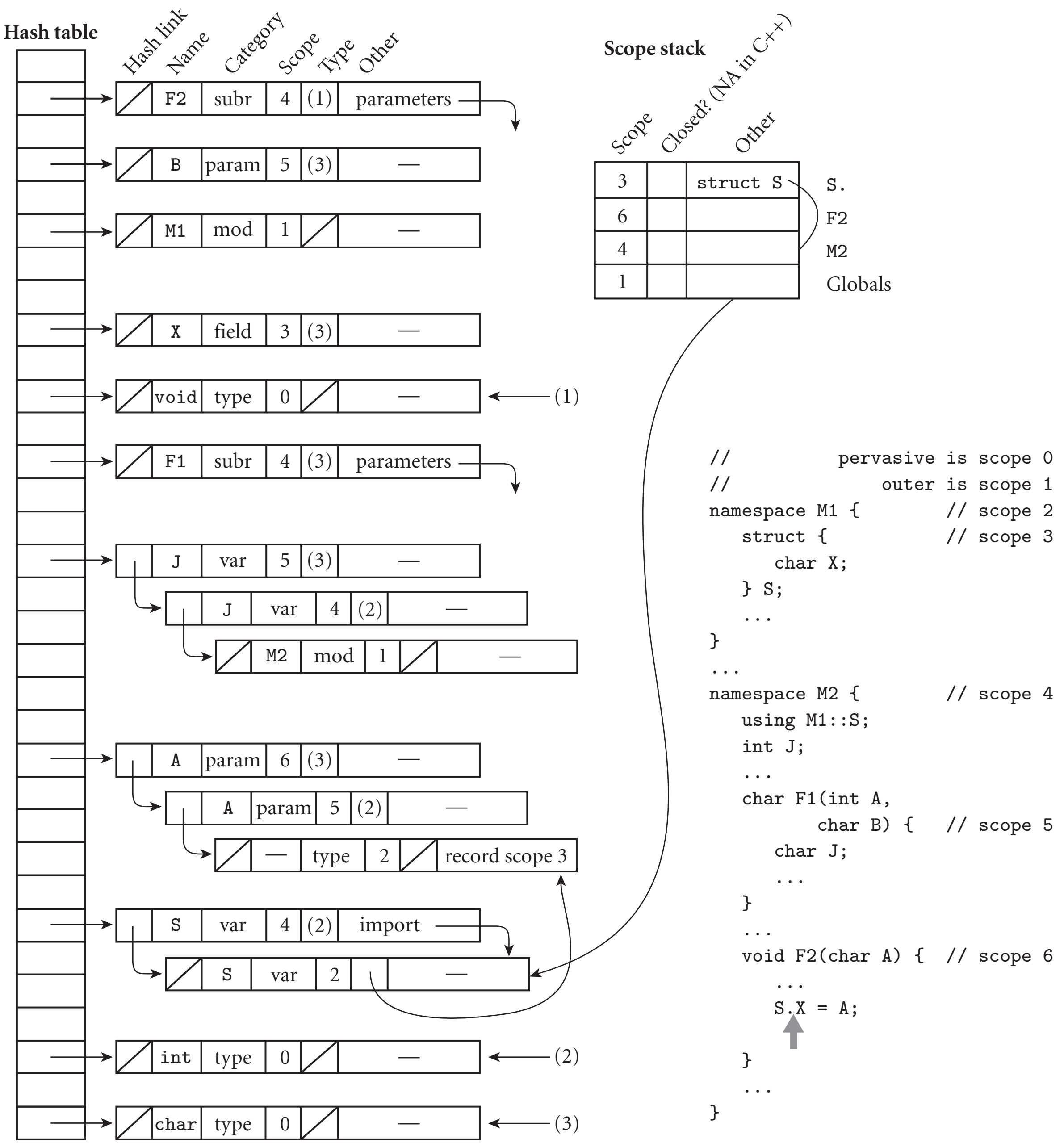
Figure 3.17 LeBlanc-Cook symbol table lookup operation.

To look up a name in the table, we scan down the appropriate hash chain looking for entries that match the name we are trying to find. For each matching entry, we scan down the scope stack to see if the scope of that entry is visible. We look no deeper in the stack than the top-most closed scope. Imports and exports are made visible outside their normal scope by creating additional entries in the table; these extra entries contain pointers to the real entries. We don't have to examine the scope stack at all for entries with scope number 0: they are pervasive. Pseudocode for the lookup algorithm appears in Figure C-3.17. ■

EXAMPLE 3.46
Symbol table for a sample program

The lower right portion of Figure C-3.18 contains the skeleton of a C++ program. The remainder of the figure shows the configuration of the symbol table for the referencing environment of the grey arrow shown in function F2. At this point in the code, the scope stack contains four entries, representing, respectively, the (anonymous) type of structure S, function F2, namespace (module) M2, and the global scope. The scope for the anonymous type indicates the specific variable (i.e., S) to which names (fields) in this scope belong. The outermost, pervasive scope is not explicitly represented.

All of the entries for a given name appear on the same hash chain, since the table is keyed on name. In this example, we assume that hash collisions have placed M2 on the same chain as the Js, and the anonymous structure type (which



```
//          pervasive is scope 0
//              outer is scope 1
namespace M1 {        // scope 2
   struct {           // scope 3
      char X;
   } S;
   ...
}
...
namespace M2 {        // scope 4
   using M1::S;
   int J;
   ...
   char F1(int A,
          char B) {   // scope 5
      char J;
      ...
   }
   ...
   void F2(char A) {  // scope 6
      ...
      S.X = A;

   }
   ...
}
```

Figure 3.18 **LeBlanc-Cook symbol table for an example program in a language like C++.** The scope stack represents the referencing environment at the grey arrow shown in function `F2`. For the sake of clarity, the many pointers from type fields to the symbol table entries for `void`, `int`, and `char` are shown as parenthesized (1)s, (2)s, and (3)s, rather than as arrows.

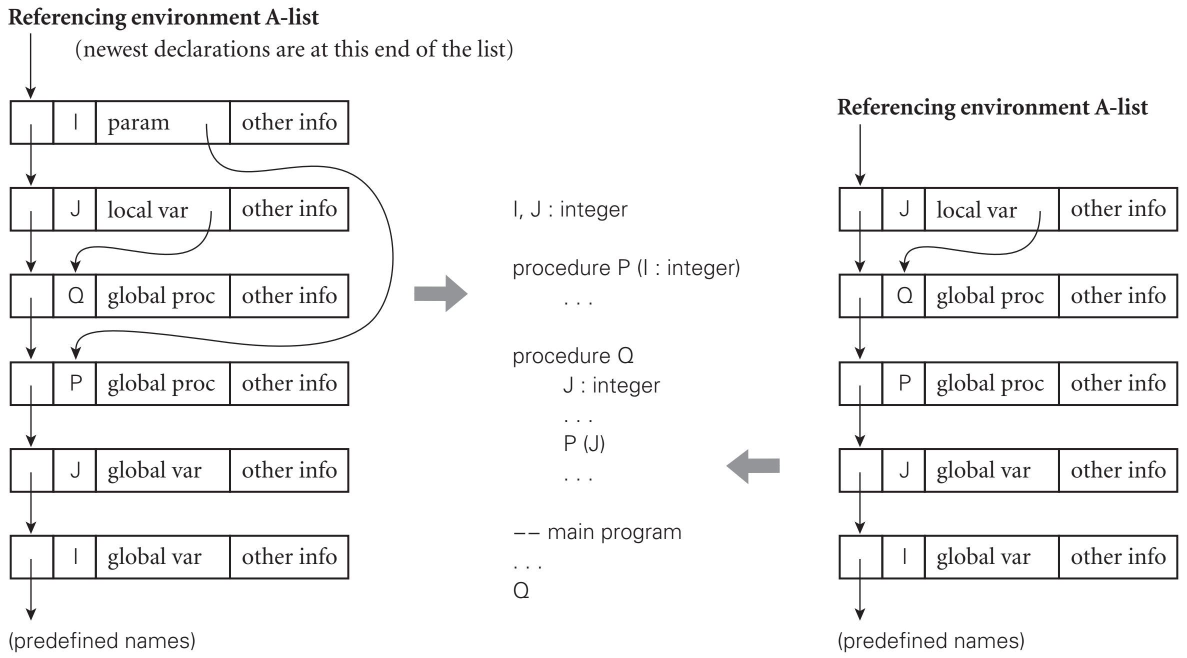
will have some arbitrary internal name) on the same chain as the `As`. Variable `S` has an extra entry, to make it directly visible inside `M2`, as requested by the `using` statement. When we are inside `F2`, a lookup operation on `J` will find `F2`'s `J`; the `J` in `M2` will be hidden by virtue of `F2` being above `M2` on the scope stack. The entry for the anonymous `struct` type indicates the scope number to be pushed onto the scope stack while resolving references to fields within objects of that type. The entry for each function contains the head pointer of a list that links together the subroutine's parameters, for use in analyzing calls (additional links of these chains are not shown). During code generation, many symbol table entries would contain additional fields, for such information as size and run-time address.

The second column of the scope stack is intended to indicate closed scopes. While C++ doesn't have any of these, we can imagine how they would work. For example, if `M2` were closed, then names in the global scope, which appears below `M2` in the scope stack, would not be visible at the indicated point in the code. Anything imported into `M2` *would* be visible, because it would have an extra entry (like that of `S`) within `M2`'s own scope.[1] If our language had exports (again, C++ does not), we would create extra entries in the *outer* scope, analogous to the ones we create in inner scopes for imports. ■

Classes, which we did not use in Figure C-3.18, would be handled much like a combination of namespaces and structures. Field and method names of the class would be visible to the class's methods, much as objects in a namespace are visible to all the namespace's code. Moreover, the entry for the class—like that of a structure type—would indicate the scope to be pushed onto the scope stack when the compiler has parsed a dot (`.`) or arrow (`->`) token and expects the next token to name a field or method of the class.

It is tempting to suggest extending a LeBlanc-Cook style symbol table to handle the visibility specifications common in object-oriented languages (e.g., the `public`, `private`, `protected` keywords of C++, to which we will return in Section 10.2.2), but this is probably a mistake. For one thing, such an extension would likely be quite messy. It is easy to make all the names in a scope visible, by pushing that scope onto the scope stack. It is also relatively easy to make a small number of names visible, by creating extra entries in other scopes, as we did for imports and exports. An intermediate option does not immediately present itself. More significantly, when the programmer attempts to use a field or method inappropriately, we probably want to generate an error message along the lines of "method m is private in class foo" instead of just "method name foo not found." This observation suggests employing a traditional lookup mechanism for class members, followed by a separate check for visibility in the current context. We consider this possibility in Exercise C-3.26.

1 Recall that the `using` statement in C++ isn't an import in the usual sense: it just gives a simpler (unqualified) name to an already-visible object.



```
I, J : integer

procedure P (I : integer)
        . . .

procedure Q
        J : integer
        . . .
        P (J)
        . . .

–– main program
. . .
Q
```

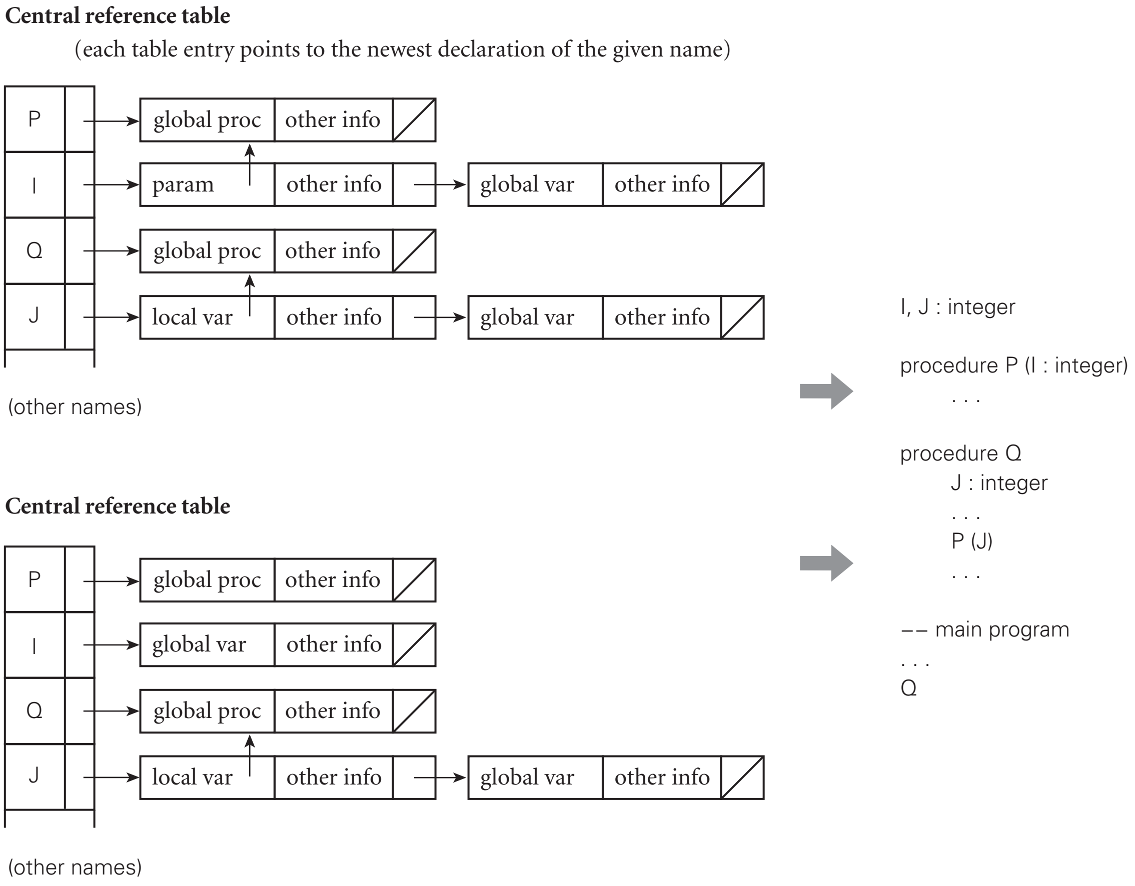
Figure 3.19 Dynamic scoping with an association list. The left side of the figure shows the referencing environment at the point in the code indicated by the adjacent grey arrow: after the main program calls Q and it in turn calls P. When searching for I, one will find the parameter at the beginning of the A-list. The right side of the figure shows the environment at the other grey arrow: after P returns to Q. When searching for I, one will find the global definition.

3.4.2 Association Lists and Central Reference Tables

Pictorial representations of the two principal implementations of dynamic scoping appear in Figures C-3.19 and C-3.20. Association lists (*A-lists*) are simple and elegant, but can be very inefficient. Central reference tables resemble a simplified LeBlanc-Cook symbol table, without the separate scope stack; they require more work at scope entry and exit than do association lists, but they make lookup operations fast.

EXAMPLE 3.47
A-list lookup in Lisp

A-lists are widely used for dictionary abstractions in Lisp; they are supported by a rich set of built-in functions in most Lisp dialects. It is therefore natural for a simple Lisp interpreter to use an A-list to keep track of name-value bindings, and even to make this list explicitly visible to the running program. Since bindings are created when entering a scope, and destroyed when leaving or returning from a scope, the A-list functions as a stack. When execution enters a scope at run time, the interpreter pushes bindings for names declared in that scope onto the top of the A-list. When execution finally leaves a scope, these bindings are removed. To look up the meaning of a name in an expression, the interpreter searches from the top of the list until it finds an appropriate binding (or reaches the end of the



```
I, J : integer

procedure P (I : integer)
        . . .

procedure Q
        J : integer
        . . .
        P (J)
        . . .

–– main program
. . .
Q
```

Figure 3.20 **Dynamic scoping with a central reference table.** The upper half of the figure shows the referencing environment at the point in the code indicated by the upper grey arrow: after the main program calls Q and it in turn calls P. When searching for I, one will find the parameter at the beginning of the chain in the I slot of the table. The lower half of the figure shows the environment at the lower grey arrow: after P returns to Q. When searching for I, one will find the global definition.

list, in which case an error has occurred). Each entry in the list contains whatever information is needed to perform semantic checks (e.g., type checking, which we will consider in Section 7.2) and to find variables and other objects that occupy memory locations. In the left half of Figure C-3.19, the first (top) entry on the A-list represents the most recently encountered declaration: the I in procedure P. The second entry represents the J in procedure Q. Below these are the global symbols, Q, P, J, and I, and (not shown explicitly) any predefined names provided by the Lisp interpreter. ■

The problem with using an association list to represent a program's referencing environment is that it can take a long time to find a particular entry in the list, particularly if it represents an object declared in a scope encountered early in the

EXAMPLE 3.48
Central reference table

program's execution, and now buried deep in the list. A central reference table is designed for faster access. It has one slot for every distinct name in the program. The table slot in turn contains a list (stack) of declarations encountered at run time, with the most recent occurrence at the beginning of the list. Looking up a name is now easy: the current meaning is found at the beginning of the list in the appropriate slot in the table. In the upper part of Figure C-3.20, the first entry on the I list is the I in procedure P; the second is the global I. If the program is compiled and the set of names is known at compile time, then each name can have a statically assigned slot in the table, which the compiled code can refer to directly. If the program is not compiled, or the set of names is not statically known, then a hash function will need to be used at run time to find the appropriate slot. ■

When control enters a new scope at run time, entries must be pushed onto the beginning of every list in the central reference table whose name is (re)declared in that scope. When control leaves a scope for the final time, these entries must be popped. The work involved is somewhat more expensive than pushing and popping an A-list, but not dramatically more so, and lookup operations are now much faster. In contrast to the symbol table of a compiler for a language with static scoping, central reference table entries for a given scope do not need to be saved when the scope completes execution; the space can be reclaimed.

Within the Lisp community, implementation of dynamic scoping via an association list is sometimes called *deep binding*, because the lookup operation may need to look arbitrarily deep in the list. Implementation via a central reference table is sometimes called *shallow binding*, because it finds the current association at the head of a given reference chain. Unfortunately, the terms "deep and shallow binding" are also more widely used for a completely different purpose, discussed in Section 3.6. To avoid potential confusion, some authors use "deep and shallow access" [Seb15] or "deep and shallow search" [Fin96] for the implementations of dynamic scoping.

Closures with Dynamic Scoping

(This subsection is best read after Section 3.6.1.)

EXAMPLE 3.49
A-list closures

If an association list is used to represent the referencing environment of a program with dynamic scoping, the referencing environment in a closure can be represented by a top-of-stack (beginning of A-list) pointer (Figure C-3.21). When a subroutine is called through a closure, the main pointer to the referencing environment A-list is temporarily replaced by the pointer from the closure, making any bindings created since the closure was created (P's I and J in the figure) temporarily invisible. New bindings created *within* the subroutine (or in any subroutine it calls) are pushed using the temporary pointer. Because the A-list is represented by pointers (rather than an array), the effect is to have two lists—one representing the caller's referencing environment and the other the temporary referencing environment resulting from use of the closure—that share their older entries. When Q returns to P in our example, the original head of the A-list will be restored, making P's I and J visible again. ■

```
procedure P(procedure C)
        declare I, J
        call C

procedure F
        declare I

procedure Q
        declare J
        call F

-- main program
        call P(Q)
```

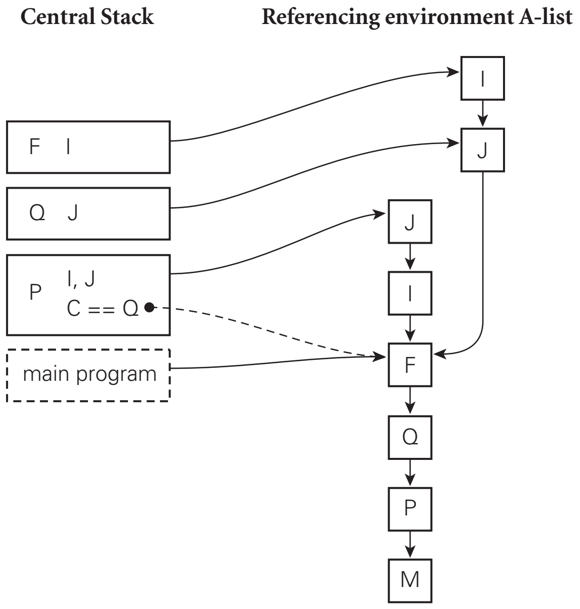


Figure 3.21 Capturing the A-list in a closure. Each frame in the stack has a pointer to the current beginning of the A-list, which the run-time system uses to look up names. When the main program passes Q to P with deep binding, it bundles its A-list pointer in Q's closure (dashed arrow). When P calls C (which is Q), it restores the bundled pointer. When Q elaborates its declaration of J (and F elaborates its declaration of I), the A-list is temporarily bifurcated.

With a central reference table implementation of dynamic scoping, the creation of a closure is more complicated. In the general case, it may be necessary to copy the entire main array of the central table and the first entry on each of its lists. Space and time overhead may be reduced if the compiler or interpreter is able to determine that only some of the program's names will be used by the subroutine in the closure (or by things that the subroutine may call). In this case, the environment can be saved by copying the first entries of the lists for only the names that will be used. When the subroutine is called through the closure, these entries can then be pushed onto the beginnings of the appropriate lists in the central reference table. Additional code must be executed to remove them again after the subroutine returns.

✓ **CHECK YOUR UNDERSTANDING**

43. List the basic operations provided by a symbol table.

44. Outline the implementation of a LeBlanc-Cook style symbol table.

45. Why don't compilers generally remove names from the symbol table at the ends of their scopes?

46. Describe the *association list* (*A-list*) and *central reference table* data structures used to implement dynamic scoping. Summarize the tradeoffs between them.
47. Explain how to implement deep binding by capturing the referencing environment A-list in a closure. Why are closures harder to build with a central reference table?

3 Names, Scopes, and Bindings

3.8 Separate Compilation

Probably the most straightforward mechanisms for separate compilation can be found in module-based languages such as Modula-2, Modula-3, and Ada, which allow a module to be divided into a declaration part (or *header*) and an implementation part (or *body*). As we noted in Section 3.3.4, the header contains all and only the information needed by users of the module (or needed by the compiler in order to compile such a user); the body contains the rest.

As a matter of software engineering practice, a design team will typically define module interfaces early in the lifetime of a project, and codify these interfaces in the form of module headers. Individual team members or subteams will then work to implement the module bodies. While doing so, they can compile their code successfully using the headers for the other modules. Using preliminary copies of the bodies, they may also be able to undertake a certain amount of testing.

In a simple implementation, only the body of a module needs to be compiled into runnable code: the compiler can read the header of module *M* when compiling the body of *M*, and also when compiling the body of any module that uses *M*. In a more sophisticated implementation, the compiler can avoid the overhead of repeatedly scanning, parsing, and analyzing *M*'s header by translating it into a symbol table, which is then accessed directly when compiling the bodies of *M* and its users. Most Ada implementations compile their module headers. Implementations of Modula-2 and 3 vary: some work one way, some the other.

As a practical matter, many languages allow the header of a module to be subdivided into a "public" part, which specifies the interface to the rest of the program, and a "private" part, which is not visible outside the module, but is needed by the compiler, for example to determine the storage requirements of opaque types. Ideally, one would include in the header of a module only that information that the programmer needs to know to use the abstraction(s) that the module provides. Restricted exports, and the public and private portions of headers, are compromises introduced to allow the compiler to generate code in the face of separate compilation.

At some point prior to execution, modules that have been separately compiled must be "glued together" to form a single program. This job is the task of the *linker*. At the very least, the linker must resolve cross-module references (loads, stores, jumps) and *relocate* any instructions whose encoding depends on the location of certain modules in the final program. Typically it also checks to make sure that users and implementors of a given interface agree on the version of the header file used to define that interface. In some environments, the linker may perform additional tasks as well, including certain kinds of interprocedural (whole-program) code improvement. We will return to the subject of linking in Chapters 15 and 16.

3.8.1 Separate Compilation in C

The initial version of C was designed at Bell Laboratories around 1970. It has evolved considerably over the years, but not, for the most part, in the area of separate compilation. Here the language remains comparatively primitive. In particular, there is in general no way for the compiler or the linker to detect inconsistencies among declarations or uses of a name in different files. The C89 standards committee introduced a new explanation of separate compilation based on the notion of *linkage*, but this served mainly to clarify semantics, not to change them. The current rules can be summarized as follows (certain details and special cases are omitted):

- If the declaration of a global object (variable or function) contains the word `static`, then the object has *internal linkage*, and is identified with (linked to) any other internally linked declaration of the same name in the same file.
- If the declaration of a function does not contain the keyword `static`, then it has *external linkage*, and is identified with any other (nonstatic) declaration of the same function in any file of the program. (A function declaration may consist of just the header.)
- If the declaration of a variable contains the keyword `extern`, then the variable has the same linkage as any visible, internally or externally linked declaration of the same name appearing earlier in the file. If there is no earlier declaration, then the variable has external linkage, and is identified with any other declaration of the same external variable in any file of the program. In other words, files in the same program that contain matching external variable declarations actually share the same variable. A global variable also has external linkage if its declaration says neither `static` nor `extern`.
- If an object is declared with both internal and external linkage, the behavior of the program is undefined.
- An object (variable or function) that is externally linked must have a *definition* in exactly one file of a program. A variable is defined when it is given an

initial value, or is declared at the global level without the `extern` keyword. A function is defined when its body (code) is given.

Many C implementations prior to C89 relaxed the final rule to permit zero or one definitions of an external variable; some permitted more than one. In these implementations, the linker unified multiple definitions, and created an implicit definition for any variable (or set of linked variables) for which the program contained only declarations.

The “linkage” rules of C89 provide a way to associate names in one file with names in another file. The rules are most easily understood in terms of their implementation. Most language-independent linkers are designed to deal with *symbols*: character-string names for locations in a machine-language program. The linker’s job is to assign every symbol a location in the final program, and to embed the address of the symbol in every machine-language instruction that makes a reference to it. To do this job, the linker needs to know which symbols can be used to resolve unbound references in other files, and which are local to a given file. C89 rules suffice to provide this information. For the programmer, however, there is no formal notion of *interface*, and no mechanism to make a name visible in some, but not all files. Moreover, nothing ensures that the declarations of an external object found in different files will be compatible: it is entirely possible, for example, to declare an external variable as a multifield record in one file and as a floating-point number in another. The compiler is not required to catch such errors, and the resulting bugs can be very difficult to find.

Header Files

Fortunately, C programmers have developed conventions on the use of external declarations that tend to minimize errors in practice. These conventions rely on the *file inclusion* facility of a macro preprocessor. The programmer creates files in pairs that correspond roughly to the interface and the implementation of a module. The name of an interface file ends with `.h`; the name of the corresponding implementation file ends with `.c`. Every object *defined* in the `.c` file is *declared* in the `.h` file. At the beginning of the `.c` file, the programmer inserts a directive that is treated as a special form of comment by the compiler, but that causes the preprocessor to include a verbatim copy of the corresponding `.h` file. This inclusion operation has the effect of placing “forward” declarations of all the module’s objects at the beginning of its implementation file. Any inconsistencies with definitions later in the file will result in error messages from the compiler. The programmer also instructs the preprocessor at the top of each `.c` file to include a copy of the `.h` files for all of the modules on which the `.c` file depends. As long as the preprocessor includes identical copies of a given `.h` file in all the `.c` files that use its module, no inconsistent declarations will occur. Unfortunately, it is easy to forget to recompile one or more `.c` files when a `.h` file is changed, and this can lead to very subtle bugs. Tools like Unix’s `make` utility help minimize such errors by keeping track of the dependences among modules.

Namespaces

Even with the convention of header files, C89 still suffers from the lack of scoping beyond the level of an individual file. In particular, all global names must be distinct, across all files of a program, and all libraries to which it links. Some coding standards encourage programmers to embed a module's name in the name of each of its external objects (e.g., `scanner_nextSym`), but this practice can be awkward, and is far from universal.

To address this limitation, C++ introduced a `namespace` mechanism that generalizes the scoping already provided for classes and functions, breaks the tie between module and compilation unit, and strengthens the interface conventions of `.h` files. Any collection of names can be declared inside a `namespace`:

EXAMPLE 3.50

Namespaces in C++

```
namespace foo {
    class foo_type_1;           // declaration
    ...
}
```

Actual definitions of the objects within `foo` can then appear in any file:

```
class foo::foo_type_1 { ...     // full definition
```

Definitions of objects declared in different namespaces can appear in the same file if desired. ■

EXAMPLE 3.51

Using names from another namespace

A C++ programmer can access the objects in a namespace using *fully qualified* names, or by *importing* (`using`) them explicitly:

```
foo::foo_type_1 my_first_obj;
```

or

```
using foo::foo_type_1;
...
foo_type_1 my_first_obj;
```

or

```
using namespace foo;    // import everything from foo
...
foo_type_1 my_first_obj;
```

There is no notion of export; all objects with external linkage in a namespace are visible elsewhere if imported or accessed with their qualified name. Note that linkage remains the foundation for separate compilation: `.h` files are merely a convention. ■

3.8.2 Packages and Automatic Header Inference

EXAMPLE 3.52
Packages in Java

The separate compilation facilities of Java and C# eliminate .h files. Java introduces a formal notion of module, called a *package*. Every *compilation unit*, which may be a file or (in some implementations) a record in a database, belongs to exactly one package, but a package may consist of many compilation units, each of which begins with an indication of the package to which it belongs:

```
package foo;
public class foo_type_1 { ...
```

Unless explicitly declared as `public`, a class in Java is visible in all and only those compilation units that belong to the same package. ■

EXAMPLE 3.53
Using names from another package

As in C++, a compilation unit that needs to use classes from another package can access them using fully qualified names, or via name-at-a-time or package-at-a-time import:

```
foo.foo_type_1 my_first_obj;
```

or

```
import foo.foo_type_1;
...
foo_type_1 my_first_obj;
```

or

```
import foo.*;           // import everything from foo
...
foo_type_1 my_first_obj;
```
■

When asked to import names from package *M*, the Java compiler will search for *M* in a standard (but implementation-dependent) set of places, and will recompile it if appropriate (i.e., if only source code is found, or if the target code is out of date). The compiler will then *automatically* extract the information that would have been needed in a C++ .h file or an Ada or Modula-3 header. If the compilation of *M* requires other packages, the compiler will search for them as well, recursively.

C# follows Java's lead in extracting header information automatically from complete class definitions. Its module-level syntax, however, is based on the namespaces of C++, which allow a single file to contain fragments of multiple namespaces. There is also no notion of standard search path in C#: to build a complete program, the programmer must provide the compiler with a complete list of all the files required.

To mimic the software engineering practice of early header file construction, a Java or C# design team can create skeleton versions of (the public classes of) its packages or namespaces, which can then be used, concurrently and independently, by the programmers responsible for the full versions.

3.8.3 Module Hierarchies

In Modula and Ada, the programmer can create a hierarchy of modules within a single compilation unit by means of lexical nesting (module `C`, for example, may be declared inside of module `B`, which in turn is declared inside of module `A`).

EXAMPLE 3.54
Multipart package names

In a similar vein, the Ada 95, Java, or C# programmer can create a hierarchy of separately compiled modules by means of *multipart names:*

```
package A.B is ...          -- Ada 95

package A.B; ...            // Java

namespace A.B { ...         // C#
```

In these examples package `A.B` is said to be a *child* of package `A`. In Ada 95 and C# the child behaves as though it had been nested inside of the parent, so that all the names in the parent are automatically visible. In Java, by contrast, multipart names work by convention only: there is no special relationship between packages `A` and `A.B`. If `A.B` needs to refer to names in `A`, then `A` must make them public, and `A.B` must import them. Child packages in Ada 95 are reminiscent of derived classes in C++, except that they support a module-as-manager style of abstraction, rather than a module-as-type style. We will consider the Ada 95 facilities further in Section 10.2.4. ■

✓ **CHECK YOUR UNDERSTANDING**

48. What purpose(s) does separate compilation serve?
49. What does it mean for an external variable to be *linked* in C?
50. Summarize the C conventions for use of `.h` and `.c` files.
51. Describe the difference between a compilation unit and a C++ or C# *namespace*.
52. Explain why Ada and similar languages separate the header of a module from its body. Explain how Java and C# get by without.

DESIGN & IMPLEMENTATION

3.12 Separate compilation

The evolution of separate compilation mechanisms from early C and Fortran, through C++, Modula-3, Ada, and finally Java and C#, reflects a move from an implementation-centric viewpoint to a more programmer-centric viewpoint. Interestingly, the ability to have zero definitions of an externally linked variable in certain early implementations of C is inherited from Fortran: the assembly language mnemonic corresponding to a declaration without a definition is `.common` (for a mechanism known as `common` blocks in Fortran).

3 Names, Scopes, and Bindings

3.10 Exercises

3.24 Assuming a LeBlanc-Cook style symbol table, explain how the compiler finds the symbol table information (e.g., the type) of a complicated reference such as `my_firm->revenues[1999]`.

3.25 Show the contents of a LeBlanc-Cook style symbol table that captures the referencing environment of

(a) function `F1` in Figure 3.4.

(b) procedure `set_seed` in Figure 3.7.

3.26 Consider the visibility of class members (fields and methods) in an object-oriented language, as discussed near the end of Section C-3.4.1. Describe a mechanism that could be used to check visibility after first locating the member in a more traditional symbol table. (You may want to look ahead to Section 10.2.2.)

3.27 Show a trace of the contents of the referencing environment A-list during execution of the program in

(a) Figure 3.9. Assume that a positive value is read at line 8.

(b) Exercise 3.14.

3.28 Repeat the previous exercise for a central reference table.

3.29 Consider the following tiny program in C:

```
void hello() {
    printf("Hello, world\n");
}

int main() {
    hello();
}
```

(a) Split the program into two separately compiled files, `tiny.c` and `hello.c`. Be sure to create a header file `hello.h` and include it correctly in `tiny.c`.

(b) Reconsider the program as C++ code. Put the `hello` function in a separate namespace, and include an appropriate `using` declaration in `tiny.c`.

(c) Rewrite the program in Java, with `main` and `hello` in separate packages.

3.30 Consider the following file from some larger C program:

```
int a;
extern int b;
static int c;

void foo() {
    int a;
    static int b;
    extern int c;
    extern int d;
}

static int b;
extern int c;
```

For each variable declaration, indicate whether the variable has external linkage, internal (file-level) linkage, or no linkage (i.e., is local).

3.31 Modula-2 provides no way to divide the header of a module into a public part and a private part: everything in the header is visible to the users of the module. Is this a major shortcoming? Are there disadvantages to the public/private division (e.g., as in Ada)? (For hints, see Section 10.2.)

3 Names, Scopes, and Bindings

3.11 Explorations

3.43 Using your favorite compiler, generate assembly language for some simple programs with debugger support enabled (on a Unix system, this will probably require the `-g` and `-S` command-line switches). Look through the result for debugger information. Can you decipher any of it? You may want to look ahead to Section 16.3.2, and to consult a manual for your system's object file format (on a modern Unix system, look for documentation on DWARF).

3.44 Learn about the *reflection* mechanisms of Java, C#, Prolog, Perl, PHP, Tcl, Python, or Ruby, all of which allow a program to inspect and reason about its own symbol table at run time. How complete are these mechanisms? (For example, can a program inspect symbols that aren't currently in scope?) What is reflection good for? What uses should be considered good or bad programming practice? For more ideas, see Section 16.3.1.

3.45 Learn about the `typeglob` mechanism of Perl, which allows a program to modify its own symbol table at run time. What are `typeglobs` good for? (See Sidebar 14.11 for some initial pointers.)

3.46 Create a C program in which a variable is exported from one file and imported by another, but the declarations in the files disagree with respect to type. You should be able to arrange for the program to compile and link successfully, but behave incorrectly. Try the same thing in Ada or C++. What happens?

3.47 Investigate the use of module hierarchies in the standard libraries of C++, Java, and C#. How is each organized? How fine grain is the division into separate headers or packages? Can you suggest an explanation for any major differences you find?

4 Semantic Analysis

4.5 Space Management for Attributes

A compiler that does not build an explicit parse tree requires some other mechanism to allocate, deallocate, and refer to storage space for attributes. In the two subsections below we consider attribute space management for bottom-up and top-down parsers, respectively. For bottom-up parsers the principal challenge is where to put the inherited attributes of symbols that have not yet been seen, and thus have no record in the parse stack. For top-down parsers this challenge does not arise, but we must go to a bit more effort to retain space for symbols that have already been parsed, and we must choose whether to manage this space automatically or to give some of the burden to the writer of action routines.

4.5.1 Bottom-Up Evaluation

EXAMPLE 4.19
Stack trace for bottom-up parse, with action routines

Figure C-4.17 shows a trace of the parse and attribute stack for `(1 + 3) * 2`, using the attribute grammar of Figure 4.1. For the sake of clarity, we show a single, combined stack for the parser and attribute evaluator, and we omit the CFSM state numbers.

It is easy to evaluate the attributes of symbols in this grammar, because the grammar is S-attributed. In an automatically generated parser, such as those produced by `yacc/bison`, the attribute rules associated with the productions of the grammar in Figure 4.1 would constitute action routines, to be executed when their productions are recognized. For `yacc/bison`, they would be written in C, with "pseudostructs" to name the attribute records of the symbols in each production. Attributes of the left-hand side symbol would be accessed as fields of the pseudostruct `$$`. Attributes of right-hand side symbols would be accessed as fields of the pseudostructs `$1`, `$2`, etc. To get from line 9 to line 10, for example, in the trace of Figure C-4.17, we would use an action routine version of the first rule of the grammar in Figure 4.1: $$.val = $1.val + $3.val. ■

When a bottom-up action routine is executed, the attribute records for symbols on the right-hand side of the production can be found in the top few entries

1. (
2. (1
3. (F_1
4. (T_1
5. (E_1
6. (E_1 +
7. (E_1 + 3
8. (E_1 + F_3
9. (E_1 + T_3
10. (E_4
11. (E_4)
12. F_4
13. T_4
14. T_4 *
15. T_4 * 2
16. T_4 * F_2
17. T_8
18. E_8

Figure 4.17 Parse/attribute stack trace for `(1 + 3) * 2`, using the grammar of Figure 4.1. Subscripts represent val attributes; they are not meant to distinguish among instances of a symbol.

of the attribute stack. The attribute record for the symbol on the left-hand side of the production (i.e., `$$`) will not yet lie in the stack: it is the task of the action routine to initialize this record. After the action routine completes, the parser pops the right-hand side records off the attribute stack and replaces them with `$$`. In `yacc/bison`, if no action routine is specified for a given production, the default action is to "copy" `$1` into `$$`. Since `$$` will occupy the same location, once pushed, that `$1` occupied before being popped, this "copy" can be effected without doing any work.

Inherited Attributes

EXAMPLE 4.20
Finding inherited attributes in "buried" records

Unfortunately, it is not always easy to write an S-attributed grammar. A simple example in which inherited attributes are desirable arises in C or Fortran-style variable declarations, in which a type name precedes the list of variable names:

$$\begin{array}{lll} \mathit{dec} & \longrightarrow & \mathit{type}\ \mathit{id\_list} \\ \mathit{id\_list} & \longrightarrow & \texttt{id} \\ \mathit{id\_list} & \longrightarrow & \mathit{id\_list}\ \texttt{,}\ \texttt{id} \end{array}$$

Let us assume that *type* has a synthesized attribute tp that contains a pointer to the symbol table entry for the type in question. Ideally, we should like to pass this attribute into *id_list* as an inherited attribute, so that we may enter each newly declared identifier into the symbol table, complete with type indication, as it is encountered. When we recognize the production *id_list* ⟶ `id`, we know that the top record on the attribute stack will be the one for `id`. But we know more than this: the next record down must be the one for *type*. To find the type of the new entry to be placed in the symbol table, we may safely inspect this "buried" record. Though it does not belong to a symbol of the current production, we can count on its presence because there is no other way to reach the *id_list* $\longrightarrow$ `id` production.

Now what about the `id` in *id_list* $\longrightarrow$ *id_list* `,` `id`? This time the top three records on the attribute stack will be for the right-hand symbols `id`, `,`, and *id_list*. Immediately below them, however, we can still count on finding the entry for *type*, waiting for the *id_list* to be completed so that *dec* can be recognized. Using nonpositive indices for pseudostructs below the current production, we can write action routines as follows:

```
dec      ⟶  type id_list
id_list  ⟶  id { declare_id ($1.name, $0.tp) }
id_list  ⟶  id_list , id { declare_id ($3.name, $0.tp) }
```

Records deeper in the attribute stack could be accessed as \$–1, \$–2, and so on. While *id_list* appears in two places in this grammar fragment, both occurrences are guaranteed to lie above a *type* record in the attribute stack, the first because it lies next to *type* in a right-hand side, and the second by induction, because it is the beginning of the yield of the first. ■

Unfortunately, there are grammars in which a symbol that needs inherited attributes occurs in productions in which the underlying symbols are not the same. We can still handle inherited attributes in such cases, but only by modifying the underlying context-free grammar.

EXAMPLE 4.21
Grammar fragment requiring context

An example can be found in languages like Perl, in which the meaning of an expression (and of the identifiers and operators within it) depends on the *context* in which that expression appears. Some Perl contexts expect arrays. Others expect numbers, strings, or Booleans. To correctly analyze an expression, we must pass the expectations of the context into the expression subtree as inherited attributes. Here is a grammar fragment that captures the problem:

```
stmt  ⟶  id := expr
      ⟶  ...
      ⟶  if expr then stmt
expr  ⟶  ...
```

Within the production for *expr*, the parser doesn't know whether the surrounding context is an assignment or the condition of an `if` statement. If it is a condition, then the expected type of the expression is Boolean. If it is an assignment, then the expected type is that of the identifier on the assignment's left-hand side. This identifier can be found two records below the current production in the attribute stack. ■

Semantic Hooks

EXAMPLE 4.22
Semantic hooks for context

To allow these cases to be treated uniformly, we can add *semantic hook*, or "marker" symbols to the grammar. Semantic hooks generate ϵ, and thus do not alter the language defined by the grammar; their only purpose is to hold inherited attributes:

$$\begin{array}{lll}
\textit{stmt} & \longrightarrow & \texttt{id := } A \ \textit{expr} \\
 & \longrightarrow & \ldots \\
 & \longrightarrow & \texttt{if } B \ \textit{expr} \ \texttt{then} \ \textit{stmt} \\
A & \longrightarrow & \epsilon \ \{ \ \$\$.\text{tp} := \${-}1.\text{tp} \ \} \\
B & \longrightarrow & \epsilon \ \{ \ \$\$.\text{tp} := \text{Boolean} \ \} \\
\textit{expr} & \longrightarrow & \ldots \{ \ \text{if } \$0.\text{tp} = \text{Boolean then} \ \ldots \}
\end{array}$$

Since the epsilon production for a semantic hook can provide an action routine, it is tempting to think of semantic hooks as a general technique to insert action routines in the middle of bottom-up productions. Unfortunately this is not the case: semantic hooks can be used only in places where the parser can be sure that it is in a given production. Placing a semantic hook anywhere else will break the "LR-ness" of the grammar, causing the parser generator to reject the modified grammar. Consider the following example:

EXAMPLE 4.23
Semantic hooks that break an LR CFG

$$\begin{array}{llll}
1. & \textit{stmt} & \longrightarrow & \textit{l\_val} \ \texttt{:=} \ \textit{expr} \\
2. & & \longrightarrow & \texttt{id} \ \textit{args} \\
3. & \textit{l\_val} & \longrightarrow & \texttt{id} \ \textit{quals} \\
4. & \textit{quals} & \longrightarrow & \textit{quals} \ \texttt{.} \ \texttt{id} \\
5. & & \longrightarrow & \textit{quals} \ \texttt{(} \ \textit{expr\_list} \ \texttt{)} \\
6. & & \longrightarrow & \epsilon \\
7. & \textit{args} & \longrightarrow & \texttt{(} \ \textit{expr\_list} \ \texttt{)} \\
8. & & \longrightarrow & \epsilon
\end{array}$$

An *l-value* in this grammar is a "qualified" identifier: an identifier followed by optional array subscript and record field qualifiers.[1] We have assumed that the language follows the notation of Fortran and Ada, in which parentheses delimit both procedure call arguments and array subscripts. In the case of procedure calls, it would be natural to want an action routine to pass the symbol-table index of the subroutine into the argument list as an inherited attribute, so that it can be used to check the number and types of arguments:

$$\begin{array}{lll}
\textit{stmt} & \longrightarrow & \texttt{id} \ A \ \textit{args} \\
A & \longrightarrow & \epsilon \ \{ \ \$\$.\text{proc\_index} := \text{lookup} \ (\$0.\text{name}) \ \}
\end{array}$$

If we try this, however, we will run into trouble, because the procedure call

1 In general, an l-value in a programming language is anything to which a value can be assigned (i.e., anything that can appear on the left-hand side of an assignment). From a low-level point of view, this is basically an address. An r-value is anything that can appear on the right-hand side of an assignment. From a low-level point of view, this is a value that can be stored at an address. We will discuss l-values and r-values further in Section 6.1.2.

```
foo(1, 2, 3);
```

and the array element assignment

```
foo(1, 2, 3) := 4;
```

begin with the same sequence of tokens. Until it sees the token after the closing parenthesis, the parser cannot tell whether it is working on production 1 or production 2. The presence of A in production 2 will therefore lead to a shift-reduce conflict; after seeing an `id`, the parser will not know whether to recognize A or shift `(`. ■

Left Corners

In general, the right-hand side of a production in a context-free grammar is said to consist of the *left corner* and the *trailing part*. In the left corner we cannot be sure which production we are parsing; in the trailing part the production is uniquely determined. In an LL(1) grammar, the left corner is always empty. In an LR(1) grammar, it can consist of up to the entire right-hand side. Semantic hooks can safely be inserted in the trailing part of a production, but not in the left corner. `Yacc/bison` recognizes this fact explicitly by allowing action routines to be embedded in right-hand sides. It automatically converts the production

EXAMPLE 4.24
Action routines in the trailing part

$S \longrightarrow \alpha$ { your code here } β

to

$S \longrightarrow \alpha\ A\ \beta$
$A \longrightarrow \epsilon$ { your code here }

for some new, distinct symbol A. If the action routine is not in the trailing part, the resulting grammar will not be LALR(1), and `yacc/bison` will produce an error message. ■

EXAMPLE 4.25
Left factoring in lieu of semantic hooks

In our procedure call and array subscript example, we cannot place a semantic hook before the *args* of production 2 because this location is in the left corner. If we wish to look up a procedure name in the symbol table before we parse the arguments, we will need to combine the productions for statements that can begin with an identifier, in a manner reminiscent of the *left factoring* discussed in Section 2.3.2:

stmt $\longrightarrow$ `id` A *quals assign_opt*
$A \longrightarrow \epsilon$ { $$.id_index := lookup ($0.name) }
quals $\longrightarrow$ *quals* `.` `id`
$\longrightarrow$ *quals* `(` *expr_list* `)`
$\longrightarrow \epsilon$
assign_opt $\longrightarrow$ `:=` *expr*
$\longrightarrow \epsilon$

This change eliminates the shift-reduce conflict, but at the expense of combining the entire grammar subtrees for procedure call arguments and array subscripts. To use the modified grammar we shall have to write action routines for *quals* that work for both kinds of constructs, and this can be a major nuisance. Users of LR-family parser generators often find that there is a tension between the desire for grammar clarity and parsability on the one hand and the need for semantic hooks to set inherited attributes on the other. ■

4.5.2 Top-Down Evaluation

Top-down parsers, as discussed in Chapter 2, come in two principal varieties: recursive descent and table driven. Attribute management in recursive descent parsers is almost trivial: inherited attributes of symbol *foo* take the form of parameters passed into the parsing routine named `foo`; synthesized attributes are the return parameters. These synthesized attributes can then be passed as inherited attributes to symbols later in the current production, or returned as synthesized attributes of the current left-hand side.

Attribute space management for automatically generated top-down parsers is somewhat more complex. Because they allow action routines at arbitrary locations in a right-hand side, top-down parsers avoid the need to modify the grammar in order to insert semantic hooks. (Of course, because they must have empty left corners, top-down grammars can be harder to write in the first place.) Because the parse stack describes the future, instead of the past, we cannot employ an attribute stack that simply mirrors the parse stack. Our two principal options are to equip the parser with a (more complicated) algorithm for automatic space management, or to require action routines to manage space explicitly.

Automatic Management

Automatic management of attribute space for top-down parsing is more complicated than it is for bottom-up parsing. It is also more space intensive. We can still use an attribute stack, but it has to contain all of the symbols in all of the productions between the root of the (hypothetical) parse tree and the current point in the parse. All of the right-hand side symbols of a given production are adjacent in the stack; the left-hand side is buried in the right-hand side of a deeper (closer to the root) production.

EXAMPLE 4.26
Operation of an LL attribute stack

Figure C-4.18 contains an LL(1) grammar for constant expressions, with action routines. Figure C-4.19 uses this grammar to trace the operation of a top-down attribute stack on the sample input `(1 + 3) * 2`. The left-hand column shows the parse stack. The right-hand column shows the attribute stack. Three global pointers index into the attribute stack. One (shown as an "arrow-boxed" `L` in the trace) identifies the record in the attribute stack that holds the attributes of the left-hand side symbol of the current production. The second (shown as an arrow-boxed `R` in the trace) identifies the first symbol on the right-hand side of the production. `L` and `R` allow the action routines to find the attributes of the

$E \longrightarrow T$ { TT.st := T.val }$^{\mathbf{1}}$ TT { E.val := TT.val }$^{\mathbf{2}}$
$TT_1 \longrightarrow$ `+` T { TT$_2$.st := TT$_1$.st + T.val }$^{\mathbf{3}}$ TT_2 { TT$_1$.val := TT$_2$.val }$^{\mathbf{4}}$
$TT_1 \longrightarrow$ `-` T { TT$_2$.st := TT$_1$.st $-$ T.val }$^{\mathbf{5}}$ TT_2 { TT$_1$.val := TT$_2$.val }$^{\mathbf{6}}$
$TT \longrightarrow \epsilon$ { TT.val := TT.st }$^{\mathbf{7}}$
$T \longrightarrow F$ { FT.st := F.val }$^{\mathbf{8}}$ FT { T.val := FT.val }$^{\mathbf{9}}$
$FT_1 \longrightarrow$ `*` F { FT$_2$.st := FT$_1$.st $\times$ F.val }$^{\mathbf{10}}$ FT_2 { FT$_1$.val := FT$_2$.val }$^{\mathbf{11}}$
$FT_1 \longrightarrow$ `/` F { FT$_2$.st := FT$_1$.st $\div$ F.val }$^{\mathbf{12}}$ FT_2 { FT$_1$.val := FT$_2$.val }$^{\mathbf{13}}$
$FT \longrightarrow \epsilon$ { FT.val := FT.st }$^{\mathbf{14}}$
$F_1 \longrightarrow$ `-` F_2 { F$_1$.val := $-$ F$_2$.val }$^{\mathbf{15}}$
$F \longrightarrow$ `(` E `)` { F.val := E.val }$^{\mathbf{16}}$
$F \longrightarrow$ `const` { F.val := C.val }$^{\mathbf{17}}$

Figure 4.18 **LL(1) grammar for constant expressions, with action routines.** The boldface superscripts are for reference in Figure C-4.19.

symbols of the current production. The third pointer (shown as an arrow-boxed `N` in the trace) identifies the first symbol within the right-hand side that has not yet been completely parsed. It allows the parser to update `L` correctly when a production is predicted.

At any given time, the attribute stack contains all symbols of all productions on the path between the root of the parse tree and the symbol currently at the top of the parse stack. Figure C-4.20 identifies these symbols graphically at the point in Figure C-4.19 immediately above the eight elided lines. Symbols to the left in the parse tree have already been reclaimed; those to the right have yet to be allocated.

At start-up, the attribute stack contains a record for the goal symbol, pointed at by `N`. When we push the right-hand side of a predicted production onto the parse stack, we add an "end-of-production" marker, represented by a colon in the trace. At the same time, we push records for the right-hand-side symbols onto the attribute stack. (These are *added* to the attribute stack; they do not replace the left-hand side.) Prior to pushing these entries, we save the current `L` and `R` pointers in another stack (not shown). We then set `L` to the old `N`, and make `R` and `N` point to the newly pushed right-hand side.

When we see an action symbol at the top of the parse stack (shown in the trace as a small bold number), we pop it and execute the corresponding action routine. When we match a terminal at the top of the parse stack, we pop it and move `N` forward one record in the attribute stack. When we see an end-of-production marker at the top of the parse stack, we pop it, set `N` to the attribute record following the one currently pointed at by `L`, pop everything from `R` forward off of the attribute stack, and restore the most recently saved values of `L` and `R`. ■

It should be emphasized that while the trace is long and tedious, its complexity is completely hidden from the writer of action routines. Once the space management routines are integrated with the driver for a top-down parser generator, all the compiler writer sees is the grammar of Figure C-4.18. In comparing Fig-

`E $`	$\boxed{N}\ E_?$
`T1TT2:$`	$\boxed{L}\ E_?\ \boxed{R}\ \boxed{N}\ T_?\ TT_{?,?}$
`F8FT9:1TT2:$`	$E_?\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ \boxed{N}\ F_?\ FT_{?,?}$
`(E) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ \boxed{N}\ (\ E_?\ )$
`E) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ (\ \boxed{N}\ E_?\ )$
`T1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ \boxed{L}\ E_?\ )\ \boxed{R}\ \boxed{N}\ T_?\ TT_{?,?}$
`F8FT9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ \boxed{N}\ F_?\ FT_{?,?}$
`C17:8FT9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ \boxed{N}\ C_1$
`17:8FT9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ C_1\ \boxed{N}$
`:8FT9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_?\ TT_{?,?}\ \boxed{L}\ F_1\ FT_{?,?}\ \boxed{R}\ C_1\ \boxed{N}$
`8FT9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ F_1\ \boxed{N}\ FT_{?,?}$
`FT9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ F_1\ \boxed{N}\ FT_{1,?}$
`14:9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_?\ TT_{?,?}\ F_1\ \boxed{L}\ FT_{1,?}\ \boxed{R}\ \boxed{N}$
`:9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_?\ TT_{?,?}\ F_1\ \boxed{L}\ FT_{1,1}\ \boxed{R}\ \boxed{N}$
`9:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ F_1\ FT_{1,1}\ \boxed{N}$
`:1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ \boxed{L}\ T_1\ TT_{?,?}\ \boxed{R}\ F_1\ FT_{1,1}\ \boxed{N}$
`1TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ \boxed{L}\ E_?\ )\ \boxed{R}\ T_1\ \boxed{N}\ TT_{?,?}$
`TT2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ \boxed{L}\ E_?\ )\ \boxed{R}\ T_1\ \boxed{N}\ TT_{1,?}$
`+T3TT4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ \boxed{L}\ TT_{1,?}\ \boxed{R}\ \boxed{N}\ +\ T_?\ TT_{?,?}$
`T3TT4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ \boxed{L}\ TT_{1,?}\ \boxed{R}\ +\ \boxed{N}\ T_?\ TT_{?,?}$
`F8FT9:3TT4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ TT_{1,?}\ +\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ \boxed{N}\ F_?\ FT_{?,?}$
`C17:8FT9:3TT4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ TT_{1,?}\ +\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ \boxed{N}\ C_3$
⟨ eight lines omitted ⟩	
`3TT4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ \boxed{L}\ TT_{1,?}\ \boxed{R}\ +\ T_3\ \boxed{N}\ TT_{?,?}$
`TT4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ \boxed{L}\ TT_{1,?}\ \boxed{R}\ +\ T_3\ \boxed{N}\ TT_{4,?}$
`7:4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ TT_{1,?}\ +\ T_3\ \boxed{L}\ TT_{4,?}\ \boxed{R}\ \boxed{N}$
`:4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ TT_{1,?}\ +\ T_3\ \boxed{L}\ TT_{4,4}\ \boxed{R}\ \boxed{N}$
`4:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ \boxed{L}\ TT_{1,?}\ \boxed{R}\ +\ T_3\ TT_{4,4}\ \boxed{N}$
`:2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ E_?\ )\ T_1\ \boxed{L}\ TT_{1,4}\ \boxed{R}\ +\ T_3\ TT_{4,4}\ \boxed{N}$
`2:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ \boxed{L}\ E_?\ )\ \boxed{R}\ T_1\ TT_{1,4}\ \boxed{N}$
`:) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_?\ FT_{?,?}\ (\ \boxed{L}\ E_4\ )\ \boxed{R}\ T_1\ TT_{1,4}\ \boxed{N}$
`) 16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ (\ E_4\ \boxed{N}\ )$
`16:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ (\ E_4\ )\ \boxed{N}$
`:8FT9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ \boxed{L}\ F_4\ FT_{?,?}\ \boxed{R}\ (\ E_4\ )\ \boxed{N}$
`8FT9:1TT2:$`	$E_?\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ F_4\ \boxed{N}\ FT_{?,?}$
`FT9:1TT2:$`	$E_?\ \boxed{L}\ T_?\ TT_{?,?}\ \boxed{R}\ F_4\ \boxed{N}\ FT_{4,?}$
`*F10FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ \boxed{L}\ FT_{4,?}\ \boxed{R}\ \boxed{N}\ *\ F_?\ FT_{?,?}$
`F10FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ \boxed{L}\ FT_{4,?}\ \boxed{R}\ *\ \boxed{N}\ F_?\ FT_{?,?}$
`C17:10FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ FT_{4,?}\ *\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ \boxed{N}\ C_2$
`17:10FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ FT_{4,?}\ *\ \boxed{L}\ F_?\ FT_{?,?}\ \boxed{R}\ C_2\ \boxed{N}$
`:10FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ FT_{4,?}\ *\ \boxed{L}\ F_2\ FT_{?,?}\ \boxed{R}\ C_2\ \boxed{N}$
`10FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ \boxed{L}\ FT_{4,?}\ \boxed{R}\ *\ F_2\ \boxed{N}\ FT_{?,?}$
`FT11:9:1TT2:$`	$E_?\ T_?\ TT_{?,?}\ F_4\ \boxed{L}\ FT_{4,?}\ \boxed{R}\ *\ F_2\ \boxed{N}\ FT_{8,?}$
⟨ six lines omitted ⟩	
`1TT2:$`	$\boxed{L}\ E_?\ \boxed{R}\ T_8\ \boxed{N}\ TT_{?,?}$
`TT2:$`	$\boxed{L}\ E_?\ \boxed{R}\ T_8\ \boxed{N}\ TT_{8,?}$
`7:2:$`	$E_?\ T_8\ \boxed{L}\ TT_{8,?}\ \boxed{R}\ \boxed{N}$
`:2:$`	$E_?\ T_8\ \boxed{L}\ TT_{8,8}\ \boxed{R}\ \boxed{N}$
`2:$`	$\boxed{L}\ E_?\ \boxed{R}\ T_8\ TT_{8,8}\ \boxed{N}$
`:$`	$\boxed{L}\ E_8\ \boxed{R}\ T_8\ TT_{8,8}\ \boxed{N}$
`$`	$E_8\ \boxed{N}$

Figure 4.19 **Trace of the parse stack (left) and attribute stack (right) for `(1 + 3) * 2`, using the grammar (and action routine numbers) of Figure C-4.18.** Subscripts in the attribute stack indicate the values of attributes. For symbols with two attributes, st comes first.

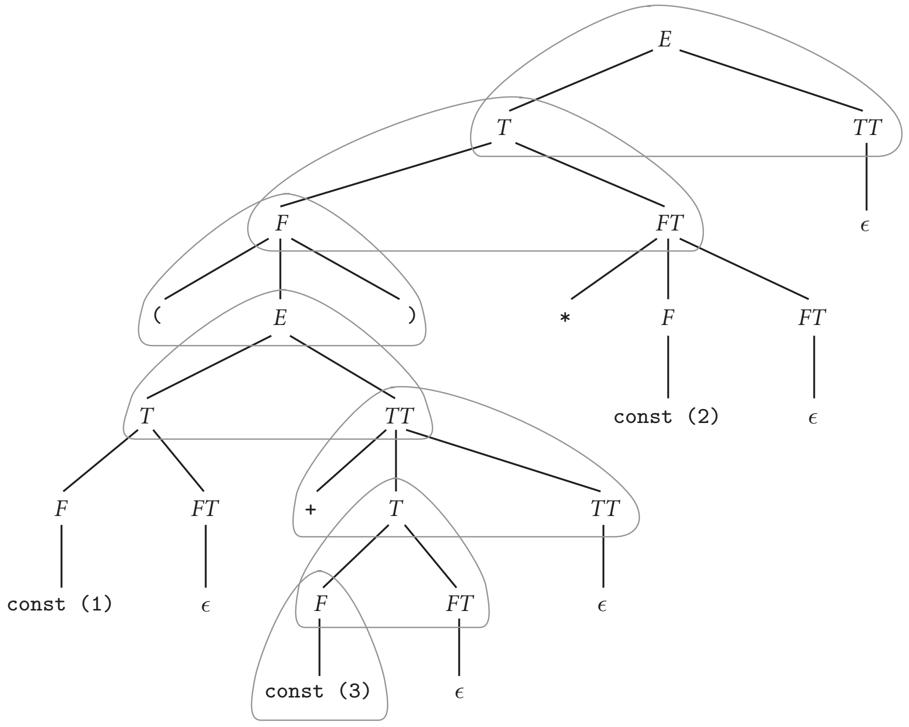


Figure 4.20 Productions with symbols currently in the attribute stack during a parse of `(1 + 3) * 2` (using the grammar of Figure C-4.18), at the point where we are about to parse the `3`. In Figure C-4.19 this point corresponds to the line immediately above the eight elided lines.

ures C-4.17 and C-4.19, one should also note that reduction and execution of a production's action routine are shown as a single step in the LR trace; they are shown separately in the LL trace, making that trace appear more complex than it really is.

Ad Hoc Management

One drawback of automatic space management for top-down grammars is the frequency with which the compiler writer must specify copy routines. Of the 17 action routines in Figure 4.9 or C-4.18, 12 simply move information from one place to another. The time required to execute these routines can be minimized by copying pointers, rather than large records, but compiler writers may still consider the copies a nuisance.

EXAMPLE 4.27
Ad hoc management of a semantic stack

An alternative is to manage space explicitly within the action routines, pushing and popping an ad hoc *semantic stack* only when information is generated or con-

$E \longrightarrow T\ TT$
$TT \longrightarrow$ `+` T { bin_op ("+") } TT
$TT \longrightarrow$ `-` T { bin_op ("−") } TT
$TT \longrightarrow \epsilon$
$T \longrightarrow F\ FT$
$FT \longrightarrow$ `*` F { bin_op ("×") } FT
$FT \longrightarrow$ `/` F { bin_op ("÷") } FT
$FT \longrightarrow \epsilon$
$F \longrightarrow$ `-` F { un_op ("$^+/_-$") }
$F \longrightarrow$ `(` E `)`
$F \longrightarrow$ `const` { push_leaf (cur_tok.val) }

Figure 4.21 Ad hoc management of attribute space in an LL(1) grammar to build a syntax tree.

sumed. Using this technique, we can replace the action routines of Figure 4.9 with the simpler version shown in Figure C-4.21. Variable cur_tok is assumed to contain the synthesized attributes of the most recently matched token. The semantic stack contains pointers to syntax tree nodes. The push_leaf routine creates a node for a specified constant and pushes a pointer to it onto the semantic stack. The un_op routine pops the top pointer off the stack, makes it the child of a newly created node for the specified unary operator, and pushes a pointer to that node back on the stack. The bin_op routine pops the top *two* pointers off the semantic stack and pushes a pointer to a newly created node for the specified binary operator. When the parse of E is completed, a pointer to a syntax tree describing its yield will be found in the top-most record on the semantic stack. ■

The advantage of ad hoc space management is clearly the smaller number of rules and the elimination of the inherited attributes used to represent left context. The disadvantage is that the compiler writer must be aware of what is in the semantic stack at all times, and must remember to push and pop it when appropriate.

One further advantage of an ad hoc semantic stack is that it allows action routines to push or pop an arbitrary number of records. With automatic space management, the number of records that can be seen by any one routine is limited by the number of symbols in the current production. The difference is particularly important in the case of productions that generate lists. In Section C-4.5.1 we saw an SLR(1) grammar for declarations in the style of C and Fortran, in which the type name precedes the list of identifiers. Here is an LL(1) grammar fragment for a language in the style of Pascal and Ada, in which the variables precede the type:

EXAMPLE 4.28
Processing lists with an attribute stack

dec $\longrightarrow$ *id_list* `:` *type*
id_list $\longrightarrow$ `id` *id_list_tail*
id_list_tail $\longrightarrow$ `,` *id_list*
$\longrightarrow \epsilon$

Without resorting to non-L-attributed flow (see Exercise C-4.28), we cannot pass the declared type into *id_list* as an inherited attribute. Instead, we must save up the list of identifiers and enter them into the symbol table *en masse* when the type is finally encountered. With automatic management of space for attributes, the action routines would look something like this:

```
dec ⟶ id_list : type { declare_vars(id_list.chain, type.tp) }
id_list ⟶ id id_list_tail { id_list.chain := append(id.name, id_list_tail.chain) }
id_list_tail ⟶ , id_list { id_list_tail.chain := id_list.chain }
             ⟶ ϵ { id_list_tail.chain := null }
```

EXAMPLE 4.29
Processing lists with a semantic stack

With ad hoc management of space, we can get by without the linked list:

```
dec ⟶ { push(marker) }
       id_list : type
       {  pop(tp)
          pop(name)
          while name ≠ marker
               declare_var(name, tp)
               pop(name)   }
id_list ⟶ id { push(cur_tok.name) } id_list_tail
id_list_tail ⟶ , id_list
             ⟶ ϵ
```

Neither automatic nor ad hoc management of attribute space in top-down parsers is clearly superior to the other. The ad hoc approach eliminates the need for many copy rules and inherited attributes, and is consequently somewhat more time and space efficient. It also allows lists to be embedded in the semantic stack. On the other hand, it requires that the programmer who writes the action routines be continually aware of what is in the stack and why, in order to push and pop it appropriately. In the final analysis, the choice is mainly a matter of taste.

✓ CHECK YOUR UNDERSTANDING

17. Explain how to manage space for synthesized attributes in a bottom-up parser.
18. Explain how to manage space for inherited attributes in a bottom-up parser.
19. Define *left corner* and *trailing part*.
20. Under what circumstances can an action routine be embedded in the right-hand side of a production in a bottom-up parser? Equivalently, under what circumstances can a marker symbol be embedded in a right-hand side without rendering the grammar non-LR?

21. Summarize the tradeoffs between automatic and ad hoc management of space for attributes in a top-down parser.
22. At any given point in a top-down parse, which symbols will have attribute records in an automatically managed attribute stack?

4 Semantic Analysis

4.8 Exercises

4.27 Repeat Exercise 4.7 using ad hoc attribute space management. Instead of accumulating the translation into a data structure, write it to a file on the fly.

4.28 Rewrite the grammar for declarations of Example C-4.28 without the requirement that your attribute flow be L-attributed. Try to make the grammar as simple and elegant as possible (you shouldn't need to accumulate lists of identifiers).

4.29 Fill in the missing lines in Figure C-4.19.

4.30 Consider the following grammar with action routines:

```
params   ⟶  mode ID par_tail
                { params.list := insert(⟨mode.val, ID.name⟩, par_tail.list) }
par_tail ⟶  , params { par_tail.list := params.list }
         ⟶  { par_tail.list := null }
mode     ⟶  IN { mode.val := IN }
         ⟶  OUT { mode.val := OUT }
         ⟶  IN OUT { mode.val := IN_OUT }
```

Suppose we are parsing the input `IN a, OUT b`, and that our compiler uses an automatically maintained attribute stack to hold the active slice of the parse tree. Show the contents of this attribute stack immediately before the parser predicts the production *par_tail* $\longrightarrow \epsilon$. Be sure to indicate where $\boxed{L}\rangle$ and $\boxed{R}\rangle$ point in the attribute stack. Also show the stack of saved $\boxed{L}\rangle$ and $\boxed{R}\rangle$ values, showing where each points in the attribute stack. You may ignore the $\boxed{N}\rangle$ pointer.

4.31 One problem with automatic space management for attributes in a top-down parser occurs in lists and sequences. Consider for example the following grammar:

$$\begin{aligned} \textit{block} &\longrightarrow \texttt{begin}\ \textit{stmt\_list}\ \texttt{end} \\ \textit{stmt\_list} &\longrightarrow \textit{stmt}\ \textit{stmt\_list\_tail} \\ \textit{stmt\_list\_tail} &\longrightarrow\ ;\ \textit{stmt\_list} \mid \epsilon \\ \textit{stmt} &\longrightarrow \ldots \end{aligned}$$

After predicting the final statement of an *n*-statement block, the attribute stack will contain the following (line breaks and indentation are for clarity only):

```
block begin stmt_list end
    stmt stmt_list_tail ; stmt_list
    stmt stmt_list_tail ; stmt_list
    stmt stmt_list_tail ; stmt_list
    { n times }
```

If the attribute stack is of finite size, it is guaranteed to overflow for some long but valid block of straight-line code. The problem is especially unfortunate since, with the exception of the accumulated output code, none of the repeated symbols in the attribute stack contains any useful attributes once its substructure has been parsed.

Suggest a technique to "squeeze out" useless symbols in the attribute stack, dynamically. Ideally, your technique should be amenable to automatic implementation, so it does not constitute a burden on the compiler writer.

Also, suppose you are using a compiler with a top-down parser that employs an automatically managed attribute stack, but does not squeeze out useless symbols. What could you do if your program caused the compiler to run out of stack space? How could you modify your program to "get around" the problem?

4 Semantic Analysis

4.9 Explorations

4.36 As described in Section C-4.5.1, `yacc/bison` will refuse to accept action routines in the left corner of a production. Is there any way around this problem? Can you imagine implementing an extended version of the tool that would permit action routines in arbitrary locations? What would be the challenges? The cost?

4.37 Learn how attribute space is managed in the ANTLR parser generator. How does it compare to the techniques described in Section C-4.5.2?

6 Control Flow

6.5.4 Generators in Icon

EXAMPLE 6.89

Simple generator in Icon

Like the iterators of Clu, Python, Ruby, and C#, Icon generators can be used for enumeration-controlled iteration. Our canonical `for` loop example would be written as follows in Icon:

```
every i := first to last by step do {
    ...
}
```

Here ... `to`... `by`... is a built-in "mixfix" generator. ■

Because Icon is intended largely for string manipulation, most of its built-in generators operate on strings. `Find(substr, str)`, for example, generates the positions (indices) within string `str` at which an occurrence of the substring `substr` can be found. `Upto(chars, str)` generates the positions within string `str` at which any character in `chars` appears. (The initial argument to `find` is a string, delimited by double quote marks; the initial argument to `upto` is a `cset` [character set], delimited by single quote marks.) The prefix operator `!` generates all elements of its operand, which can be a string, list, record, file, or table.

EXAMPLE 6.90

A generator inside an expression

In comparison to conventional iterators, however, the generators of Icon are more deeply embedded in the semantics of the language. A generator can be used in any context that expects an expression. The larger context is then capable of generating multiple results. The following code will print all positions in `s` that follow a blank:

```
every i := 1 + upto(' ', s) do {
    write(i)
}
```

This can even be written as

```
every write(1 + upto(' ', s))
```

■

EXAMPLE 6.91

Generating in search of success

Generators in Icon are used not only for iteration, but also for *goal-directed search*, implemented via *backtracking*. (Backtracking is also fundamental to Prolog, which we will study in Chapter 12.) Where most languages use Boolean expressions to control selection and logically controlled loops, Icon uses a more general notion of *success* and *failure*. A conditional statement such as

```
if 2 < 3 then {
    ...
}
```

is said to execute not because the condition `2 < 3` is true, but because the *comparison* `2 < 3` *succeeds*. The distinction is important for generators, which are capable of producing results repeatedly until one of them causes the surrounding context to succeed (or until no more results can be produced). For example, in

```
if (i := find("abc", s)) > 6 then {
    ...
}
```

the body of the `if` statement will be executed only if the string `"abc"` appears beyond the sixth position in `s`. Because `find` generates its results in order, `i` will represent the first such position (if any). The execution model is as follows: `find` is capable of generating all positions at which `"abc"` occurs in `s`. Suppose the first such occurrence is at position 2. Then `i` is assigned the value 2, but the comparison `2 > 6` fails. Because there is a generator inside the failed expression, Icon will resume that generator and reevaluate the expression for the next generated value. It will continue this reevaluation process until the comparison succeeds, or until the generator runs out of values, in which case it (the generator) fails, the overall expression fails definitively, and the body of the `if` is skipped. ■

EXAMPLE 6.92

Backtracking with multiple generators

If a failed expression contains more than one generator, all possible values will be explored systematically. The body of the following `if`, for example, will be executed if and only if an `x` appears at the same position in both `s` and `t`, with `i` denoting the first such matching position:

```
if (i := find("x", s)) = find("x", t) then {
    ...
}
```

If there is no matching position, then `i` will be set to the position of the final `x` in `s`, but the body of the loop will be skipped. If the programmer wishes to avoid changing `i` in the case where the overall test fails, then the *reversible assignment* operator, `<-` can be used instead of `:=`. When Icon backtracks past a reversible assignment, it restores the original value. ■

Any user-defined subroutine in Icon can be a generator if it uses the `suspend` *expr* statement instead of `return` *expr*. `Suspend` is Icon's equivalent of `yield`. If the expression following `suspend` contains an invocation of a generator, then the subroutine will suspend repeatedly, once for each generated value.

✓ CHECK YOUR UNDERSTANDING

45. Explain how Icon generators differ from the iterators of Clu, Python, Ruby, and C#, and from the iterator objects of Euclid, C++, and Java.
46. Describe the notions of *success* and *failure* in Icon.
47. What is *backtracking*? Why is it useful?
48. Name a language other than Icon in which backtracking plays a fundamental role.

6 Control Flow

6.7 Nondeterminacy

In Algol 68, the lack of ordering among expression operands was explicitly defined as an example of nondeterminacy, which the language designers called *collateral execution*. Several other built-in constructs in Algol 68 were nondeterministic as well, and an explicit *collateral statement* allowed the programmer to specify nondeterminacy in the evaluation of arbitrary expressions when desired.

Among his many contributions to the art of programming, Dijkstra [Dij75] advocated the use of nondeterminacy for selection and logically controlled loops. His *guarded command* notation has been adopted by several languages. One of these is SR, a pedagogical language of the 1980s, which we will mention again in Chapter 13. Imagine for a moment that we are writing a function to return the maximum of two integers. In C, we would probably employ a code fragment something like this:

EXAMPLE 6.93 Avoiding asymmetry with nondeterminism

```
if (a > b) max = a;
else max = b;
```

Of course, we could also write

```
if (a >= b) max = a;
else max = b;
```

These fragments differ in their behavior when a = b: the first sets `max = b`; the second sets `max = a`. As a practical matter the difference is irrelevant, since a and b are equal, but it is in some sense aesthetically unpleasant to have to make an arbitrary choice between the two. More important, the arbitrariness of the choice makes it more difficult to reason about the code formally, or to prove it is correct. In a language with guarded commands (the example here is in SR), one could write the following:

EXAMPLE 6.94 Selection with guarded commands

```
if a >= b -> max := a
[] b >= a -> max := b
fi
```

The general form of this construct is

```
if condition -> stmt_list
[] condition -> stmt_list
[] condition -> stmt_list
...
fi
```

■

Each of the conditions in this construct is known as a *guard*. The guard and its following statement, together, are called a *guarded command*. When control reaches an `if` statement in a language with guarded commands, a nondeterministic choice is made among the guards that evaluate to true, and the statement list following the chosen guard is executed. In SR, the final condition may optionally be `else`. If none of the conditions evaluates to true, the statement list following the `else`, if any, is executed. If there is no `else`, the `if` statement as a whole has no effect. (In Dijkstra's original proposal, there was no `else` guard option, and it was a dynamic semantic error for none of the guards to be true.) Interestingly, SR provides no separate `case` construct: the SR compiler detects when the conditions of an `if` statement test the same expression against a nonoverlapping set of compile-time constants, and generates table-lookup code as appropriate.

EXAMPLE 6.95

Looping with guarded commands

SR uses guarded commands for several purposes in addition to selection. Its logically controlled looping construct (again patterned on Dijkstra's proposal) looks very much like the `if` statement:

```
do condition -> stmt_list
[] condition -> stmt_list
[] condition -> stmt_list
...
od
```

For each iteration of the loop, a nondeterministic choice is made among the guards that evaluate to true, and the statement list following the chosen one is executed. The loop terminates when none of the guards is true (there is no `else` guard option for loops). Using this notation, we can write Euclid's greatest common divisor algorithm as follows:

```
do a > b -> a := a - b
[] b > a -> b := b - a
od
gcd := a
```

■

```
process client:
    loop
        toss coin
        if heads, send read request to server
                  wait for response
        if tails, send write request to server
                  wait for response

process server:
    loop
            receive read request
            reply with data
        OR
            receive write request
            update data and reply
```

Figure 6.7 **Example of a concurrent program that requires nondeterminacy.** The server must be able to accept either a `read` or a `write` request, whichever is available at the moment. If it insists on receiving them in any particular order, deadlock may result.

Nondeterministic Concurrency

EXAMPLE 6.96
Nondeterministic message receipt

While nondeterministic constructs have a certain appeal from an aesthetic and formal semantics point of view, their most compelling advantages arise in concurrent programs, for which they can affect correctness. Imagine, for example, that we are writing a simple dictionary program to support computer-aided design on a network of personal computers. The dictionary keeps a mapping from part names to their specifications. A dictionary server process handles requests from clients on other workstations on the network. Each request may be either a read (return me the current specification for part X) or a write (define part Y as follows).[1] Clients send requests at unpredictable times. As a result, the server cannot tell at any given time whether it should try to receive a read or a write request. If it makes the wrong choice the entire system may deadlock (see Figure C-6.7). ■

EXAMPLE 6.97
Nondeterministic server in SR

Most message-based concurrent languages provide at least one mechanism to specify nondeterministic choice among potential communication partners. These mechanisms do not all look like guarded commands, but they have similar semantics. In SR, one could write our dictionary server as follows:

```
# declarations of request types:
op read_data(n : name) returns d : description
op write_data(n : name; d : description)
```

1 This is of course an oversimplified example. Among other things, any real system of this sort would need a mechanism to lock parts in the dictionary, so that no two clients would ever end up designing new specifications for the same part concurrently.

```
# local subroutines:
proc lookup ...         # find info in dictionary
proc update ...         # change info in dictionary
# code for server:
process server
    do true ->          # loop forever
        in read_data(n) returns d -> d := lookup(n)
        [] write_data(n, d) -> update(n, d)
        ni
    od
end
```

Here `in` is a nondeterministic construct whose guards can contain communication statements. The guard `write_data(n, d)` will evaluate to true if and only if some client is attempting to send a request containing a new specification for a part. We shall see in Section C-13.5.3 that more elaborate guards can allow a server to constrain the types of requests that it is willing to receive at a given point in time, or even to "peek" inside a message to see if it is acceptable. If none of the guards of an `in` statement is true, the server waits until one is. ■

Choosing among Guards

EXAMPLE 6.98
Naive (unfair) implementation of nondeterminism

What happens if two or more guards evaluate to true? How does the language implementation choose among them? We have glossed over this issue so far. The most naive implementation would treat a guarded command construct like a conventional `if`... `then` ... `else`:

```
server:
    loop
        if read_data request available
            ...
        elsif write_data request available
            ...
        else wait until some request is available
```

The problem with this implementation is that it always favors one type of request over another; if `read_data` requests are always available, `write_data` requests will never be received. ■

EXAMPLE 6.99
"Gotcha" in round-robin implementation of nondeterminism

A slightly more sophisticated implementation would maintain a circular list of the guards in each set of guarded commands. Each time it encounters the construct in which these commands appear, it would check guards beginning with the one after the one that succeeded last time. This technique works well in many cases, but can fail consistently in others. In the following, for example (again in SR), the guard of the first `in` statement combines a communication test with a Boolean condition:

```
process silly
var count : int := 0
    do true ->
        in A() st count % 2 = 1 -> ...
        [] B() -> ...
        [] C() -> ...
        ni
        count++
    od
```

This example is somewhat contrived, but illustrates the problem. The `st` ("such that") clause in the first guard indicates that it can be chosen only on odd iterations of the loop. Now imagine that `A`, `B`, and `C` requests are always available. If we always check guards starting with the one after the one that succeeded last time (beginning at first with the initial guard), then `B` will be chosen in the first iteration (because `count mod 2` $\neq 1$), `C` will be chosen in the second iteration (when `count` $= 2$), `B` will be chosen again in the third iteration (because again `count mod 2` $\neq 1$), and so forth. `A` will never be chosen. The lesson to be learned from this example is that no deterministic algorithm will provide a truly satisfactory implementation of a nondeterministic construct (see Sidebar C-6.11). ■

One final issue has to do with side effects. Guarded command constructs make a nondeterministic choice among the guards that evaluate to true. They do *not*, however, guarantee that all guards will be evaluated before the choice is made; the implementation is free to ignore the rest of the guards once it has chosen one that is true. A program may therefore produce unexpected or even unpredictable

DESIGN & IMPLEMENTATION

6.11 Nondeterminacy and fairness

Ideally, what we should like in a nondeterministic construct is a guarantee of *fairness*. This turns out to be trickier than one might expect: there are several plausible ways that "fair" might be defined. Certainly we should like to guarantee that no guard that is always true is always skipped. Probably, we should like to guarantee that no guard that is true infinitely often (in a hypothetical infinite sequence of iterations) is always skipped. Better, we might ask that any guard that is true infinitely often be chosen infinitely often. This stronger notion of fairness will obtain if the choice among true guards is genuinely random. Unfortunately, good pseudorandom number generators are expensive enough that we may not want to use them to choose among guards. As a result, most implementations of guarded commands are not provably fair. Many simply employ the circular list technique. Others use somewhat "more random" heuristics. Many machines, for example, provide a fast-running clock register that can be read efficiently in user-level code. A reasonable "random" choice of the guard to evaluate first can be made by interpreting this clock as an integer, and computing its remainder modulo the number of guards.

results if any of the guards have side effects. This problem is the programmer's responsibility in SR. An alternative would have been to prohibit side effects and have the compiler verify their absence.

✓ CHECK YOUR UNDERSTANDING

49. What is a *guarded command*?
50. Explain why nondeterminacy is particularly important for concurrent programs.
51. Give three alternative definitions of *fairness* in the context of nondeterminacy.
52. Describe three possible ways of implementing the choice among guards that evaluate to true. What are the tradeoffs among these?

6 Control Flow

6.9 Exercises

6.36 (David Hanson [Han93].) Write a program in Icon that will print the *k* most common words in its input, one per line, with each preceded by a count of the number of times it appears. If parameter *k* is not specified on the command line, use 10 by default. You will want to consult the Icon manual (available on-line [GG96]). In addition to `suspend`, `upto`, and `write`, discussed in this text, you may find it helpful to learn about `integer`, `many`, `pull`, `read`, `sort`, `table`, and `tab`. When fed the Gettysburg Address, your program should print

```
13  that
9   the
8   we
8   to
8   here
7   a
6   and
5   of
5   nation
5   have
```

6.37 Write a `findRE` generator in Icon that mimics the behavior of `find`, but takes as its first parameter a regular expression. Use a string to represent your regular expression, with syntax as in Section 2.1.1. Use empty parentheses to represent ϵ. Give highest precedence to Kleene closure, then concatenation, then alternation. You may assume that we never search for vertical bar, asterisk, or parenthesis characters.

6.38 Explain why the following guarded commands in SR are *not* equivalent:

```
if a < b -> c := a              if a < b -> c := a
[] b < c -> c := b              [] b < c -> c := b
[] else -> c := d               [] true -> c := d
fi                              fi
```

6.39 The astute reader may have noticed that the final line of the code in Example C-6.95 embodies an arbitrary choice. It could just as easily have said `gcd := b`. Show how to use a guarded command to restore the symmetry of the program.

6.40 Write, in SR or pseudocode, a function that returns

(a) an arbitrary nonzero element of a given array

(b) an arbitrary permutation of a given array

In each case, write your code in such a way that if the implementation of nondeterminism were truly random, all correct answers would be equally likely.

6 Control Flow

6.10 Explorations

6.47 Learn about Snobol, an earlier language by Ralph Griswold, who also designed Icon. How do the two languages compare?

6.48 Chapter 18 of Griswold's text on Icon [GG96] discusses scanning and parsing. After reading this chapter, explain how backtracking search can be used to generalize recursive descent. What classes of grammars can you parse with this generalized technique? What is the worst-case time complexity?

6.49 Learn about the `select` routine in the Unix (POSIX) library. How does it deal with the need for nondeterministic receipt from multiple communication partners? How would you use this routine to achieve the effect of the SR code in Example C-6.97?

6.50 Explain how to use threads in Java to achieve the effect of Example C-6.97.

7 Type Systems

7.3.2 Generics in C++, Java, and C#

Though templates were not officially added to C++ until 1990, when the language was almost ten years old, they were envisioned early in its evolution. C# generics, likewise, were planned from the beginning, though they actually didn't appear until the 2.0 release in 2004. By contrast, generics were deliberately omitted from the original version of Java. They were added to Java 5 (also in 2004) in response to strong demand from the user community.

C++ Templates

EXAMPLE 7.58
Generic `arbiter` class in C++

Figure C-7.5 defines a simple generic class in C++ that we have named an `arbiter`. The purpose of an `arbiter` object is to remember the "best instance" it has seen of some generic parameter class `T`. We have also defined a generic `chooser` class that provides an `operator()` method, allowing it to be called like a function. The intent is that the second generic parameter to `arbiter` should be a subclass of `chooser`, though this is not enforced. Given these definitions we might write

```
class case_sensitive : chooser<string> {
public:
    bool operator()(const string& a, const string& b) { return a < b; }
};
...
arbiter<string, case_sensitive> cs_names;       // declare new arbiter
cs_names.consider(new string("Apple"));
cs_names.consider(new string("aardvark"));
cout << *cs_names.best() << "\n";               // prints "Apple"
```

Alternatively, we might define a `case_insensitive` descendant of `chooser`, whereupon we could write

```
template<typename T>
class chooser {
public:
    virtual bool operator()(const T& a, const T& b) = 0;
};

template<typename T, typename C>
class arbiter {
    T* best_so_far;
    C comp;
public:
    arbiter() { best_so_far = nullptr; }
    void consider(T* t) {
        if (!best_so_far || comp(*t, *best_so_far)) best_so_far = t;
    }
    T* best() {
        return best_so_far;
    }
};
```

Figure 7.5 Generic arbiter in C++.

```
arbiter<string, case_insensitive> ci_names;     // declare new arbiter
ci_names.consider(new string("Apple"));
ci_names.consider(new string("aardvark"));
cout << *ci_names.best() << "\n";               // prints "aardvark"
```

Either way, the C++ compiler will create a new instance of the arbiter template every time we declare an object (e.g., cs_names) with a different set of generic arguments. Only when we attempt to use such an object (e.g., by calling consider) will it check to see whether the arguments support all the required operations.

Because type checking is delayed until the point of use, there is nothing magic about the chooser class. If we neglected to define it, and then left it out of the header of case_sensitive (and similarly case_insensitive), the code would still compile and run just fine. ■

C++ templates are an extremely powerful facility. Template parameters can include not only types, but also values of ordinary (nongeneric) types, and nested template declarations. Programmers can also define *specialized* templates that provide alternative implementations for certain combinations of arguments. These facilities suffice to implement recursion, giving programmers the ability, at least in principle, to compute arbitrary functions at compile time (in other words, templates are *Turing complete*). An entire branch of software engineering has grown up around so-called *template metaprogramming*, in which templates are used to persuade the C++ compiler to generate custom algorithms for special circumstances [AG05]. As a comparatively simple example, one can write a template that accepts a generic parameter int n and produces a sorting routine for *n*-element arrays in which all of the loops have been completely unrolled.

As described in Section 7.3.1 ("Implicit Instantiation"), C++ allows generic parameters to be *inferred* for generic functions, rather than specified explicitly. To identify the right version of a generic function (from among an arbitrary number of specializations), and to deduce the corresponding generic arguments, the compiler must perform a complicated, potentially recursive pattern-matching operation. This pattern matching is, in fact, quite similar to the type inference of ML-family languages, described in Section 7.2.4. It can, as noted in Sidebar 7.8, be cast as *unification*.

Unfortunately, per-use instantiation of templates has several significant drawbacks. First, it requires that the compiler have access to the template's source code at the point in the program where instantiation occurs. In the code of Figure C-7.5, the `arbiter` class includes complete definitions of its methods. This is entirely appropriate for small, simple classes, even in a header (`.h`) file. If the code were significantly more complex, we might wish to put only the declaration of the generic class in our header file (call it `arbiter.h`), and defer the method definitions to a separate `arbiter.cc` file:

EXAMPLE 7.59
Template function bodies moved to a `.c` file

```
// arbiter.h:

template<typename T, typename C>
class arbiter {
    T* best_so_far;
    C comp;
public:
    arbiter();
    void consider(T* t);
    T* best();
};

// arbiter.cc (imagine that these methods were long and complicated):

template<class T, class C>
arbiter<T,C>::arbiter() { best_so_far = nullptr; }

template<class T, class C>
void arbiter<T,C>::consider(T* t) {
    if (!best_so_far || comp(*t, *best_so_far)) best_so_far = t;
}

template<class T, class C>
T* arbiter<T,C>::best() { return best_so_far; }
```

Compilation units that have access to the `.h` file will still compile successfully, but now the actual code of the `arbiter` methods will never be instantiated. The likely symptom will be "missing symbol" errors from the linker.

C++ provides a partial solution to this problem, in the form of *explicit instantiation*. If we anticipate the need for case-sensitive and case-insensitive `string` arbiters, we can define the appropriate `chooser` classes in `arbiter.h`, and then instantiate corresponding `arbiter` classes in `arbiter.cc`:

```
template class arbiter<string, case_sensitive>;
template class arbiter<string, case_insensitive>;
```

■

Of course, explicit instantiation works only if the implementor of a template's `.cc` file knows what instantiations will eventually be required. If this cannot be anticipated, the bodies will need to remain in the `.h` file, regardless of their complexity. But then a second problem arises: if the same template is instantiated with the same arguments in 20 different compilation units, the compiler will end up compiling the same code 20 times. Most modern linkers are smart enough to keep only one copy of the machine code for a repeatedly instantiated template, but we will have wasted not only the cost of repeated scanning and parsing, but of semantic analysis, optimization, and code generation as well.

EXAMPLE 7.60
`extern` templates in C++11

C++11 provides a partial solution to this second problem, in the form of `extern` template declarations. If the templated class declaration and method definitions of Example C-7.59 were included in their entirety in `arbiter.h`, and we then needed a case-sensitive `arbiter` in each of 20 `.cc` files, we could write

```
extern template class arbiter<string, case_sensitive>;
```

in all but one of the files, instructing the compiler *not* to generate machine code for that `arbiter`, but rather to assume that an appropriate implementation would be generated elsewhere (presumably in the 20th file, where the `extern` keyword would be omitted), and would thus be available at link time. ■

EXAMPLE 7.61
Instantiation-time errors in C++ templates

In day-to-day use, the final and perhaps the most frustrating problem with per-use instantiation is its tendency to result in inscrutable error messages. Continuing our running example, if we define

```
class foo {                                  // line 31 of source
public:
    bool operator()(const string& a, const unsigned int b) {
        // wrong type for second parameter, from arbiter's point of view
        return a.length() < b;
    }
};
```

and then say

```
arbiter<string, foo> oops;
...
oops.consider(new string("Apple"));          // line 66 of source
```

one might hope to receive an error message along the lines of "line 66: foo's operator() method needs to take two arguments of type string&." Instead the GNU C++ compiler responds

```
best.cc: In instantiation of 'void arbiter<T, C>::consider(T*)
[with T = std::basic_string<char>; C = foo]':
best.cc:66:38:   required from here
best.cc:19:26: error: no match for call to
'(foo) (std::basic_string<char>&, std::basic_string<char>&)'
         if (!best_so_far || comp(*t, *best_so_far)) best_so_far = t;
                          ^
best.cc:31:7: note: candidate is:
 class foo {
       ^
best.cc:33:10: note: bool foo::operator()(const string&, unsigned int)
     bool operator()(const string& a, const unsigned int b) {
          ^
best.cc:33:10: note:   no known conversion for argument 2
    from 'std::basic_string<char>' to 'unsigned int'
```

LLVM's `clang` front end is similarly inscrutable:

```
best.cc:19:29: error: no matching function for call to
object of type 'foo'
        if (!best_so_far || comp(*t, *best_so_far)) best_so_far = t;
                            ^~~~
best.cc:66:10: note: in instantiation of member function
'arbiter<std::__1::basic_string<char>, foo>::consider' requested here
    oops.consider(new string("Apple"));
         ^
best.cc:33:10: note: candidate function not viable: no known conversion
from 'std::__1::basic_string<char>' to 'const unsigned int'
for 2nd argument
    bool operator()(const string& a, const unsigned int b) {
         ^
```

The problem here is fundamental; it's not poor compiler design. Because the language requires that templates be "expanded out" before they are type checked, it is extraordinarily difficult to generate messages without reflecting that expansion. One of the principal goals of the generic parameter constraints currently under consideration for the next release of C++ is to improve the quality of error messages, by performing more high-level checks before fully expanding the template. ■

Java Generics

Generics were deliberately omitted from the original version of Java. Rather than instantiate containers with different generic parameter types, Java programmers followed a convention in which all objects in a container were assumed to be of the standard base class `Object`, from which all other classes are descended. Users of a container could place any type of object inside. When removing an object,

```
interface Chooser<T> {
    public boolean better(T a, T b);
}

class Arbiter<T> {
    T bestSoFar;
    Chooser<? super T> comp;

    public Arbiter(Chooser<? super T> c) {
        comp = c;
    }
    public void consider(T t) {
        if (bestSoFar == null || comp.better(t, bestSoFar)) bestSoFar = t;
    }
    public T best() {
        return bestSoFar;
    }
}
```

Figure 7.6 Generic arbiter in Java.

a *cast* could be needed to reassert the original type. No danger was involved, because objects in Java are self-descriptive, and casts employ run-time checks.

Though dramatically simpler than the use of templates in C++, this programming convention has three significant drawbacks: (1) users of containers must litter their code with casts, which many people find distracting or aesthetically distasteful; (2) errors in the use of a container manifest themselves as `ClassCastExceptions` at run time, rather than as compile-time error messages; (3) the casts incur overhead at run time. Given Java's emphasis on clarity of expression, rather than pure performance, problems (1) and (2) were considered the most serious, and became the subject of a Java Community Process proposal for a language extension in Java 5. The solution adopted is based on the GJ (Generic Java) work of Bracha et al. [BOSW98].

EXAMPLE 7.62
Generic `Arbiter` class in Java

Figure C-7.6 contains a Java version of our `arbiter` class. It differs from the C++ code of Figure C-7.5 in several important ways. First, Java requires that the code for each generic class be manifestly (self-obviously) type safe, independent of any particular instantiation. This means that the type of field `comp`—and in particular, the fact that it provides a `better` method—must be statically declared. As a result, the `Chooser` to be used by a given `Arbiter` instance must be specified as a constructor parameter; it cannot be a generic parameter. (We could have used a constructor parameter in C++; in Java it is mandatory.) For both field `comp` and constructor parameter `c`, we are then faced with the question: what should be the generic parameter of `Chooser`?

The most obvious choice (*not* the one adopted in Figure C-7.6) would be `Chooser<T>`. This would allow us to write

```
class CaseSensitive implements Chooser<String> {
    public boolean better(String a, String b) {
        return a.compareTo(b) < 1;
    }
}
...
Arbiter<String> csNames = new Arbiter<String>(new CaseSensitive());
csNames.consider(new String("Apple"));
csNames.consider(new String("aardvark"));
System.out.println(csNames.best());           // prints "Apple"
```
■

EXAMPLE 7.63
Wildcards and bounds on Java generic parameters

Suppose, however, we were to define

```
class CaseInsensitive implements Chooser<Object> {   // note type!
    public boolean better(Object a, Object b) {
        return a.toString().compareToIgnoreCase(b.toString()) < 1;
    }
}
```

Class `Object` defines a `toString` method (usually used for debugging purposes), so this declaration is valid. Moreover since every `String` is an `Object`, we ought to be able to pass any pair of strings to `CaseInsensitive.better` and get a valid response. Unfortunately, `Chooser<Object>` is not acceptable as a match for `Chooser<String>`. If we typed

```
Arbiter<String> ciNames = new Arbiter<String>(new CaseInsensitive());
```

the compiler would complain. The fix (as shown in Figure C-7.6) is to declare both `comp` and `c` to be of type `<? super T>` instead. This informs the Java compiler that an arbitrary type argument ("`?`") is acceptable as the generic parameter of our `Chooser`, so long as that type is an ancestor of `T`.

The `super` keyword specifies a *lower bound* on a type parameter. It is the symmetric opposite of the `extends` keyword, which we used in Example 7.51 to specify an *upper bound*. Together, upper and lower bounds allow us to broaden the set of types that can be used to instantiate generics. As a general rule, we use `extends T` whenever a method returns a `T` object (on which we need to be able to invoke `T` methods); we use `super T` whenever we expect to pass a `T` object as a parameter, but don't mind if the receiver is willing to accept something more general. Given the bounded declarations of Figure C-7.6, our use of `CaseInsensitive` will compile and run just fine:

```
Arbiter<String> ciNames = new Arbiter<String>(new CaseInsensitive());
ciNames.consider(new String("Apple"));
ciNames.consider(new String("aardvark"));
System.out.println(ciNames.best());           // prints "aardvark"
```
■

```
interface Chooser {
    public boolean better(Object a, Object b);
}

class Arbiter {
    Object bestSoFar;
    Chooser comp;

    public Arbiter(Chooser c) {
        comp = c;
    }
    public void consider(Object t) {
        if (bestSoFar == null || comp.better(t, bestSoFar)) bestSoFar = t;
    }
    public Object best() {
        return bestSoFar;
    }
}
```

Figure 7.7 **Arbiter in Java after type erasure.** No casts are required in this portion of the code (but see the main text for uses).

Type Erasure

EXAMPLE 7.64
Type erasure and implicit casts

Generics in Java are defined in terms of *type erasure*: the compiler effectively deletes every generic parameter and argument list, replaces every occurrence of a type parameter with `Object`, and inserts casts back to concrete types wherever objects are returned from generic methods. The erased equivalent of Figure C-7.6 appears in Figure C-7.7. No casts are required in this portion of the code. On any use of `best`, however, the compiler would insert an implicit cast. The statement

```
String winner = csNames.best();
```

will, in effect, be implicitly replaced with

```
String winner = (String) csNames.best();
```

Also, in order to match the `Chooser<String>` interface, our definition of `CaseSensitive` (Example C-7.62) will in effect be replaced with

```
class CaseSensitive implements Chooser {
    public boolean better(Object a, Object b) {
        return ((String) a).compareTo((String) b) < 1;
    }
}
```

The advantage of type erasure over the nongeneric (`Object`-based) version of the code is that the programmer doesn't have to write the casts. In addition, the compiler is able to verify in most cases that the erased code will never generate a `ClassCastException` at run time. The exceptions occur primarily when, for the sake of interoperability with preexisting code, the programmer assigns a generic collection into a nongeneric collection:

EXAMPLE 7.65
`Unchecked` warnings in Java

```
Arbiter<String> csNames = new Arbiter<String>(new CaseSensitive());
Arbiter alias = csNames;            // nongeneric
alias.consider(new Integer(3));     // unsafe
```

The compiler will issue an "unchecked" warning on the third line of this example, because we have invoked method `consider` on a "raw" (nongeneric) `Arbiter` without explicitly casting the arguments. In this case the warning is clearly warranted: `alias` *shouldn't* be passed an `Integer`. Other examples can be quite a bit more subtle. It should be emphasized that the warning simply indicates the lack of *static* checking; any type errors that actually occur will still be caught at run time. ■

EXAMPLE 7.66
Java generics and built-in types

Note, by the way, that the use of erasure, and the insistence that every instance of a given generic be able to share the same code, means that type arguments in Java must all be descended from `Object`. While `Arbiter<Integer>` is a perfectly acceptable type, `Arbiter<int>` is not. ■

DESIGN & IMPLEMENTATION

7.11 Why erasure?

Erasure in Java has several surprising consequences. For one, we can't invoke `new T()`, where `T` is a type parameter: the compiler wouldn't know what kind of object to create. Similarly, Java's *reflection* mechanism, which allows a program to examine and reason about the concrete type of an object at run time, knows nothing about generics: `csNames.getClass().toString()` returns `"class Arbiter"`, not `"class Arbiter<String>"`. Why would the Java designers introduce a mechanism with such significant limitations? The answer is backward compatibility or, more precisely, *migration* compatibility, which requires complete interoperability of old and new code.

More so than most previous languages, Java encourages the assembly of working programs, often on the fly, from components written independently by many different people in many different organizations. The Java designers felt it was critical not only that old (nongeneric) programs be able to run with new (generic) libraries, but also that new (generic) programs be able to run with old (nongeneric) libraries. In addition, they took the position that the Java virtual machine, which interprets Java bytecode in the typical implementation, could not be modified. While one can take issue with these goals, once they are accepted erasure becomes a natural solution.

C# Generics

Though generics were omitted from C# version 1, the language designers always intended to add them, and the .NET Common Language Infrastructure (CLI) was designed from the outset to provide appropriate support. As a result, C# 2.0 was able to employ an implementation based on *reification* rather than erasure. Reification creates a different concrete type every time a generic is instantiated with different arguments. Reified types are visible to the reflection library (`csNames.GetType().ToString()` returns `"Arbiter`1[System.Double]"`), and it is perfectly acceptable to call `new T()` if `T` is a type parameter with a zero-argument constructor (a constraint to this effect is required). Moreover where the Java compiler must generate implicit type casts to satisfy the requirements of the virtual machine (which knows nothing of generics) and to ensure type-safe interaction with legacy code (which might pass a parameter or return a result of an inappropriate type), the C# compiler can be sure that such checks will never be needed, and can therefore leave them out. The result is faster code.

EXAMPLE 7.67
Sharing generic implementations in C#

Of course the C# compiler is free to merge the implementations of any generic instantiations whose code would be the same. Such sharing is significantly easier in C# than it is in C++, because implementations typically employ just-in-time compilation, which delays the generation of machine code until immediately prior to execution, when it's clear whether an identical instantiation already exists somewhere else in the program. In particular, `MyType<Foo>` and `MyType<Bar>` will share code whenever `Foo` and `Bar` are both classes, because C# employs a reference model for variables of class type. ■

EXAMPLE 7.68
C# generics and built-in types

Like C++, C# allows generic arguments to be value types (built-ins or `structs`), not just classes. We are free to create an object of class `MyType<int>`; we do not have to "wrap" it as `MyType<Integer>`, the way we would in Java. `MyType<int>` and `MyType<double>` would generally not share code, but both would run significantly faster than `MyType<Integer>` or `MyType<Double>`, because they wouldn't incur the dynamic memory allocation required to create a wrapper object, the garbage collection required to reclaim it, or the indirection overhead required to access the data inside. ■

Like Java, C# allows only types as generic parameters, and insists that generics be manifestly type safe, independent of any particular instantiation. It generates reasonable error messages if we try to instantiate a generic with an argument that doesn't meet the constraints of the corresponding generic parameter, or if we try, inside the generic, to invoke a method that the constraints don't guarantee will be available.

EXAMPLE 7.69
Generic `Arbiter` class in C#

A C# version of our `Arbiter` class appears in Figure C-7.8. One small difference with respect to Figure C-7.6 can be seen in the `Arbiter` constructor, which must explicitly initialize field `bestSoFar` to `default(T)`. We can leave this out in Java because variables of class type are implicitly initialized to `null`, and type parameters in Java are all classes. In C# `T` might be a built-in or a `struct`, both of which require explicit initialization. ■

```
interface Chooser<in T> {
    bool better(T a, T b);
}

class Arbiter<T> {
    T bestSoFar;
    Chooser<T> comp;
    bool initialized;

    public Arbiter(Chooser<T> c) {
        comp = c;
        bestSoFar = default(T);
        initialized = false;
    }
    public void Consider(T t) {
        if (!initialized || comp.better(t, bestSoFar)) bestSoFar = t;
        initialized = true;
    }
    public T Best() {
        return bestSoFar;
    }
}
```

Figure 7.8 Generic arbiter in C#.

EXAMPLE 7.70

Contravariance in the `Arbiter` interface

A more interesting difference from Figure C-7.6 appears in the definitions of the `Chooser` interface, the `comp` member of class `Arbiter`, and the `c` parameter of the `Arbiter` constructor. In Java, we used explicit lower bounds (`? super T`) on `comp` and `c` to indicate that any `Chooser<S>`, where `S` is a superclass of `T`, would be acceptable. While C# allows us to specify upper bounds in the form of type constraints (we did so in the `sort` routine of Example 7.52), it has no direct equivalent of lower bounds. It does, however, support the related notions of *covariance* and *contravariance*. We have exploited this support in Figure C-7.8, where it appears not as bounds on the `Chooser` passed to a newly created `Arbiter`, but as an `in` modifier on the generic parameter of the `Chooser` interface itself.

The declaration `interface Chooser<in T>` indicates that objects of class `T` will be used only as input parameters to methods of the interface. Suppose now that `S` is a superclass of `T`. Since `T` provides all the methods of `S`, any method that expects an input of class `S` will also accept an input of class `T`. This means that in any context in which all we do is provide `T` objects as inputs to a `Chooser`, we can use a "less choosy" `Chooser` that merely expects `S` inputs. In other words, `Chooser<T>` is a superclass of `Chooser<S>`. Represented graphically,

$$\texttt{T} \rightarrow \texttt{S} \;\Rightarrow\; \texttt{Chooser<S>} \rightarrow \texttt{Chooser<T>}$$

where the $\rightarrow$ symbol, pronounced "is a," indicates that the item on the left inherits from the item on the right. `Chooser<T>` is said to be "*contra*variant in `T`" because the relationship between `S` and `T` is reversed when wrapping them in a `Chooser`. ■

EXAMPLE 7.71
Covariance

In other situations, objects of a generic type may only be *produced* by the methods of an interface. Consider, for example, the notion of an iterator, as provided by C#'s `IEnumerator<T>` interface. Method `Current` of this interface returns an object of class `T`; no method takes a `T` object as input. In the C# standard library, the interface is declared as

```
public interface IEnumerator<out T> ...
```

Now suppose again that `S` is a superclass of `T`. In any context in which all we do is extract `S` objects from an `IEnumerator`, we can use a more specific `IEnumerator` that gives us `T` objects instead. In other words, `IEnumerator<S>` is a superclass of `IEnumerator<T>`. Graphically,

$$\texttt{T} \rightarrow \texttt{S} \;\Rightarrow\; \texttt{IEnumerator<T>} \rightarrow \texttt{IEnumerator<S>}$$

Here `IEnumerator<T>` is said to be "*co*variant in `T`" because the relationship between `S` and `T` is preserved when wrapping them in an `IEnumerator`. In many interfaces, of course, generic parameters appear as both inputs and outputs of methods. For such an interface `Foo`, there is no subclassing relationship: `Foo<T>` is said to be "*in*variant in `T`." ■

EXAMPLE 7.72
`Chooser` as a delegate

Returning to the `Arbiter` example, there is actually a simpler way to write our code in C#. Because the `Chooser` interface has only a single method, we can express it as a *delegate* instead:

```
delegate bool Chooser<T>(T a, T b);
```

Then in method `Arbiter.Consider`, we can call the delegate directly as `comp(t, bestSoFar)`. Our new `Chooser` is roughly analogous to the C declaration

```
typedef _Bool (*Chooser)(T a, T b);
```

(pointer to function of two `T` arguments, returning a Boolean), except that a C# `Chooser` object is a closure, not a pointer: it can refer to a static function, a method of a particular object (in which case it has access to the object's fields), or an anonymous nested function (in which case it has access, with unlimited extent, to variables in the surrounding scope). In our particular case, defining `Chooser` to be a delegate allows us to pass any appropriate function to the `Arbiter` constructor, without regard to the class inheritance hierarchy. We can declare

```
static bool CaseSensitive(String a, String b) {
    return String.CompareOrdinal(a, b) < 1;
    // use Unicode order, in which upper-case letters come first
}
static bool CaseInsensitive(Object a, Object b) {
    return String.Compare(a.ToString(), b.ToString(), false) < 1;
}
```

and then say

```
Arbiter<String> csNames =
    new Arbiter<String>(new Chooser<String>(CaseSensitive));
csNames.Consider("Apple");
csNames.Consider("aardvark");
Console.WriteLine(csNames.Best());              // prints "Apple"

Arbiter<String> ciNames =
    new Arbiter<String>(new Chooser<String>(CaseInsensitive));
ciNames.Consider("Apple");
ciNames.Consider("aardvark");
Console.WriteLine(ciNames.Best());              // prints "aardvark"
```

The compiler is perfectly happy to instantiate `CaseInsensitive` as a `Chooser<String>`, because `String`s can be passed as `Object`s. ■

✓ CHECK YOUR UNDERSTANDING

36. Why is it difficult to produce high-quality error messages for misuses of C++ templates?
37. What is the purpose of explicit instantiation in C++? What is the purpose of `extern` templates?
38. What is *template metaprogramming*?
39. Explain the difference between *upper bounds* and *lower bounds* in Java type constraints. Which of these does C# support?
40. What is *type erasure*? Why is it used in Java?
41. Under what circumstances will a Java compiler issue an “unchecked” generic warning?
42. Why must fields of generic parameter type be explicitly initialized in C#?
43. For what two main reasons are C# generics often more efficient than comparable code in Java?
44. Summarize the notions of *covariance* and *contravariance* in generic types.
45. How does a C# *delegate* differ from an interface with a single method (e.g., the C++ `chooser` of Figure C-7.5)? How does it differ from a function pointer in C?

7 Type Systems

7.6 Exercises

7.21 C++ has no direct analogue of the `extends X` and `super X` clauses of Java. Why not?

7.22 Write a simple abstract `ordered_set<T>` class (an *interface*) whose methods include `void insert(T val)`, `void remove (T val)`, `bool lookup (T val)`, and `bool is_empty()`, together with a language-appropriate iterator, as described in Section 6.5.3. Using this abstract class as a base, build a simple `list_set` class that uses a sorted linked list internally. Try this exercise in C++, Java, and C#. Note that (in Java and C#, at least) you will need constraints on `T`. Discuss the differences among your implementations.

7.23 Building on the previous exercise, implement higher-level `union<T>`, `intersection<T>`, and `difference<T>` functions that operate on ordered sets. Note that these should not be members of the `ordered_set<T>` class, but rather stand-alone functions: they should be independent of the details of `list_set` or any other particular `ordered_set`. So, for example, `union(A, B, C)` should verify that `A` is empty, and then add to it all the elements found in `B` or `C`. Explain, for each of C++, Java, and C#, how to handle the comparison of elements.

7.24 Continuing Example C-7.62, the call

```
csNames.consider(null);
```

will generate a run-time exception, because `String.compareTo` is not designed to take `null` arguments.

(a) Modify Figure C-7.6 to guard against this possibility by including a predicate `public Boolean valid(T a);` in the `Chooser<T>` interface, and by modifying `consider` to make an appropriate call to this predicate. Modify class `CaseSensitive` accordingly.

(b) Suggest how to make similar modifications to the C# `Arbiter` of Figure C-7.8 and Example C-7.69. How should you handle lower bounds when you need both `Better` and `Valid`?

7.25 (a) Modify your solution to Exercise 7.14 so that the comparison routine is an explicit generic parameter, reminiscent of the `chooser` of Figure C-7.5.

(b) Give an alternative solution in which the comparison routine is an extra parameter to `sort`.

7.26 Consider the C++ program shown in Figure C-7.9. Explain why the final call to `first_n` generates a compile-time error, but the call to `last_n` does not. (Note that `first_n` is generic but `last_n` is not.) Show how to modify the final call to `first_n` so that the compiler will accept it.

7.27 Consider the following code in C++:

```
template <typename T>
class cloneable_list : public list<T> {
public:
    cloneable_list<T>* clone() {
        auto rtn = new cloneable_list<T>();
        for (auto e : *this) {
            rtn->push_back(e);
        }
        return rtn;
    }
};

...
cloneable_list<foo> L;
...
cloneable_list<foo>* Lp = L.clone();
```

Here `*Lp` will be a "deep copy" of `L`, containing a copy of each `foo` object. Try to write equivalent code in Java. What goes wrong? How might you get around the problem?

```
#include <iostream>
#include <list>
using std::cout;
using std::list;

template<typename T> void first_n(list<T> p, int n) {
    for (typename list<T>::iterator li = p.begin(); li != p.end(); li++) {
        if (n-- <= 0) break;
        cout << *li << " ";
    }
    cout << "\n";
}

void last_n(list<int> p, int n) {
    for (list<int>::reverse_iterator li = p.rbegin(); li != p.rend(); li++) {
        if (n-- <= 0) break;
        cout << *li << " ";
    }
    cout << "\n";
}

class int_list_box {
    list<int> content;
public:
    int_list_box(list<int> l) { content = l; }
    operator list<int>() { return content; }
        // user-supplied operator for coercion/conversion
};

int main() {
    int i = 5;
    list<int> l;

    for (int i = 0; i < 10; i++) l.push_back(i);
    int_list_box b(l);

    first_n(l, i);      // works
    last_n(b, i);       // works (coerces b)
    first_n(b, i);      // static semantic error
}
```

Figure 7.9 Coercion and generics in C++. The compiler refuses to accept the final call to `first_n`.

7 Type Systems

7.7 Explorations

7.36 Learn about the *Concepts Lite* proposal for generic parameter constraints in C++ [SSD13], support for which is already available in experimental versions of `gcc`. Compare and contrast this proposal with the constraint mechanisms of Java and C#.

7.37 Explore the support for generics in Scala, Eiffel, Ada, or some other programming language. Compare this support to that of C++, Java, and C#. What might account for the differences? Which approach(es) do you prefer? Why?

7.38 Explore more fully the concepts of *covariance* and *contravariance* in object-oriented languages, as exemplified by the `in` and `out` modifiers for generic parameters in C# 4.0. Discuss the connection between these concepts and the notions of upper and lower bounds on generic parameters (`? extends T` and `? super T` in Java).

8 Composite Types

8.1.3 Variant Records (Unions)

A variant record provides two or more alternative fields or collections of fields, only one of which is valid at any given time. This notion has its roots in the `equivalence` statement of Fortran I and in the `union` types of Algol 68. Building on the `element` type of Example 8.1, one could implement a variant record as follows in (pre-2011) C:

EXAMPLE 8.59
Nested `struct`s and `union`s in traditional C

```
struct element {
    char name[2];
    int atomic_number;
    double atomic_weight;
    _Bool metallic;
    _Bool naturally_occurring;
    union {
        struct {
            char *source;
                /* textual description of principal commercial source */
            double prevalence;
                /* fraction, by weight, of Earth's crust */
        } natural_info;
        double lifetime;
            /* half-life in seconds of most stable known isotope */
    } extra_fields;
} copper;
```

Here the programmer presumably intends for the `naturally_occurring` field to indicate which parts of the union are currently valid. A `true` value indicates that the element has at least one naturally occurring stable isotope; in this case fields `source` and `prevalence` are intended to describe how the element may be obtained and how commonly it occurs. A `false` value indicates that the element results only from atomic collisions or the decay of heavier elements; in this case, field `lifetime` is intended to indicate how long atoms so created tend to survive before undergoing radioactive decay. These mutually exclusive sets of fields

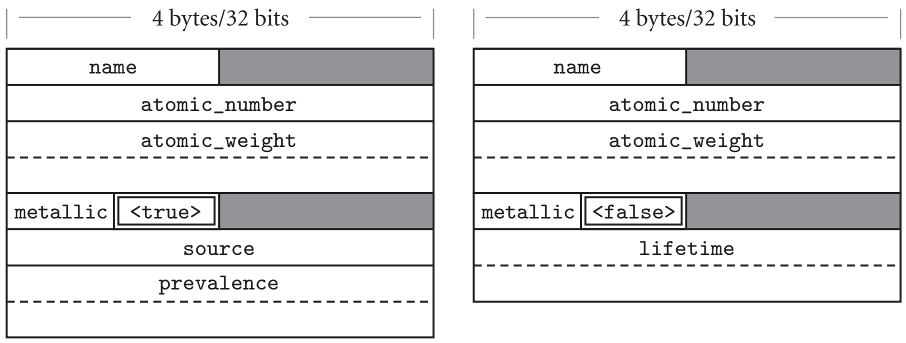


Figure 8.16 **Likely memory layouts for `element` variants.** The value of the `naturally_occurring` field (shown here with a double border) is intended to indicate which of the interpretations of the remaining space is valid. Field `source` is assumed to point to a string that has been independently allocated.

(`source` and `prevalence`, on the one hand, or `lifetime` on the other) are sometimes known as *variants*. Either the first or the second variant may be useful, but never both at once. From an implementation perspective, nonoverlapping uses suggest that the variants may share space, as shown in Figure C-8.16. ■

One significant problem with our nested `struct` and `union` is the need for two extra levels of naming. While the always-present fields can be accessed as, say, `copper.atomic_weight`, fields of the inner `struct` are much less easy to name: `copper.extra_fields.natural_info.source`.

EXAMPLE 8.60
A variant record in Pascal

Pascal's principal contribution to union types was to integrate them with records. In Pascal syntax, our running example might look like this:

```
type element = record
    name : two_chars;
    atomic_number : integer;
    atomic_weight : real;
    metallic : Boolean;
    case naturally_occurring : Boolean of
      true : (
        source : string_ptr;
        prevalence : real;
      );
      false : (
        lifetime : real;
      )
end;
```

Here the `naturally_occurring` field is introduced with the keyword `case`, to formalize its role as a *tag* or *discriminant*. Note that the variant fields have no extraneous levels of naming: we can refer directly to `copper.source`. ■

EXAMPLE 8.61

Anonymous unions in C11 and C++11

Leveraging an extension long supported by gcc, C11 allows a nameless (*anonymous*) struct or union to appear within another struct or union. The members of the anonymous construct are then directly visible in the surrounding context:

```
struct element {
    char name[2];
    int atomic_number;
    double atomic_weight;
    _Bool metallic;
    _Bool naturally_occurring;
    union {
        struct {
            char *source;
            double prevalence;
        };
        double lifetime;
    };
} copper;
...
copper.source = "various ores";
```

Anonymous nesting makes variants in C11 as convenient as those of Pascal. C++11 even allows anonymous unions in non-struct contexts:

```
void foo() {
    union {
        int a;
        int b;
    };
    ...
    a = 3;
```

■

Safety

EXAMPLE 8.62

Breaking type safety with unions

A potentially more significant problem with unions in C is the lack of type safety. Mistakes in which the programmer writes to one field of a union and then reads from the other are relatively common:

```
union {
    int i;
    double d;
} u;
...
u.d = 3.0;
...
printf("%d", u.i);
```

Here the `printf` statement, which attempts to output `i` as an integer, will (in most implementations) take its bits from the floating-point representation of

3.0—almost certainly a mistake, but one that the language implementation will not catch. ■

To avoid these sorts of errors, Algol 68 included features to track the status of unions at run time, and to prevent access to currently invalid fields. Similar features can be found in Ada and in ML-family languages today. Our running `element` example might be written as follows in OCaml:

EXAMPLE 8.63
Type-safe unions in OCaml

```
type natural_info = {source : string; prevalence : float};;
type synthesized_info = {lifetime : float};;
type extra_info =
    Natural of natural_info
  | Synthesized of synthesized_info;;

type element = {
    name : string;
    atomic_number : int;
    atomic_weight : float;
    metallic : bool;
    extra_fields : extra_info};;
```

As in traditional C, the variant portions of a record introduce extra levels of nesting in OCaml. To enforce correct usage, the language implementation maintains a hidden tag in every union object, to indicate which variant is currently valid. Values can be declared only as aggregates that specify the tag and all the fields:

```
let copper = {
    name = "Cu";
    atomic_number = 29;
    atomic_weight = 63.546;
    metallic = true;
    extra_fields = Natural ({
        source = "various ores and native deposits";
        prevalence = 0.00005
    })
};;
```

Individual fields can be read, but only in the context of a `match` expression that verifies the value of the tag:

```
exception Union_error;;
let source (e : element) =
    match e.extra_fields with
      | Natural     n -> n.source
      | Synthesized _ -> raise Union_error;;

let copper_source = source copper;;
```

■

Variant records with mandatory tags (explicit or hidden) are known as *discriminated unions*. Variant records without tags (as in C) are known as *nondiscriminated unions*. Pascal provided both, but in the absence of an analogue of `match`, even the discriminated case was difficult to implement safely (more on this in Exercise C-8.26). Ada, by contrast, combines syntax reminiscent of Pascal with the type safety of ML.

Variants in Ada

EXAMPLE 8.64
Ada variants and tags (discriminants)

Ada variant records must always have a tag (called the *discriminant*). Language rules ensure that this tag can never be changed without simultaneously assigning values to all of the fields of the corresponding variant. The assignment can occur either via whole-record assignment (e.g., `a := b`, where `a` and `b` are variant records), or via assignment of an aggregate (e.g., `p := (polar => true, rho => 1.0, theta => pi/2.0))`. In addition to appearing as a field within the record, the discriminant of a variant record in Ada must also appear in the header of the record's declaration:

```
type element (naturally_occurring : Boolean := true) is record
    name : string (1..2);
    atomic_number : integer;
    atomic_weight : long_float;
    metallic : Boolean;
    case naturally_occurring is
        when true =>
            source : string_ptr;
            prevalence : long_float;
        when false =>
            lifetime : long_float;
    end case;
end record;
```

Here we have not only declared the discriminant of the record in its header, we have also specified a default value for it. A declaration of a variable of type `element` has the option of accepting this default value:

```
copper : element;
```

or overriding it:

```
plutonium : element (false);
neptunium : element (naturally_occurring => false);
    -- alternative syntax
```

If the type declaration for `element` did not specify a default value for `naturally_occurring`, then all variables of type `element` would have to provide a value. These rules guarantee that the tag field of a variant record is never uninitialized. ■

An Ada record variable whose declaration specifies a value for the discriminant is said to be *constrained*. Its tag field can never be changed by a subsequent assignment. This immutability means that the compiler can allocate just enough space to hold the specified variant; this space may in some cases be significantly smaller than would be required for other variants. A variable whose declaration does not provide an initial value for the discriminant is said to be *unconstrained*. Its tag will be initialized to the value in the type declaration, but may be changed by later (whole-record) assignments, so the space that the record occupies must be large enough to hold any possible variant.

EXAMPLE 8.65
A discriminated subtype in Ada

An Ada subtype definition can also constrain the discriminant(s) of its parent type:

```
subtype natural_element is element (true);
```

Variables of type `natural_element` will all be constrained; their `naturally_occurring` field cannot be changed. Because `natural_element` is a subtype, rather than a derived type, values of type `element` and `natural_element` are compatible with each other, though a run-time semantic check will usually be required to assign the former into the latter. ■

EXAMPLE 8.66
Discriminated array in Ada

Ada uses record discriminants not only for variant tags, but in general for any value that affects the size of a record. Here is an example that uses a discriminant to specify the length of an array:

DESIGN & IMPLEMENTATION

8.13 The placement of variant fields

To facilitate space saving in constrained variant records, Ada requires that all variant parts of a record appear at the end. This rule ensures that every field has a constant offset from the beginning of the record, with no holes (in any variant) other than those required for alignment. When a constrained variant record is elaborated, the Ada run-time system need only allocate sufficient space to hold the specified variant, which is never allowed to change. Pascal had a similar rule, designed for a similar purpose. When a variant record was allocated from the heap in Pascal (via the built-in `new` operator), the programmer had the option of specifying `case` labels for the variant portions of the record. A record so allocated was never allowed to change to a different variant, so the implementation could allocate precisely the right amount of space.

Modula-2, which did not provide `new` as a built-in operation, eliminated the ordering restriction on variants. All variables of a variant record type had to be large enough to hold any variant. The usual implementation assigned a fixed offset to every field, with holes following small internal variants as necessary. Similar conventions apply to `union`s and `struct`s in modern C.

```
type element_array is array (integer range <>) of element;
type alloy (num_components : integer) is record
    name : string (1..30);
    components : element_array (1..num_components);
    tensile_strength : long_float;
end record;
```

The <> notation in the initial definition of element_array indicates that the bounds are not statically known. Further discussion of dynamic arrays appears in Section 8.2.2. As with discriminants used for variant tags, the programmer must either specify a default value for the discriminant in the type declaration (we did not do so above), or else every declaration of a variable of the type must specify a value for the discriminant (in which case the variable is constrained, and the discriminant cannot be changed). ■

The Object-Oriented Alternative

In dropping variant records from their parent language, the designers of Modula-3 noted [Har92, p. 110] that much of the same effect could be obtained with classes and inheritance. Oberon, similarly, replaced variants with a more general mechanism for *type extension* (Section 10.2.4), and the designers of Java and C# dropped the unions of C and C++. In place of the C code of Example C-8.59, a Java programmer might write

EXAMPLE 8.67
Derived types as an alternative to unions

```
class Element {
    public String name;
    public int atomic_number;
    public double atomic_weight;
    public boolean metallic;
}
class NaturalElement extends Element {
    public String source;
    public double prevalence;
}
class SyntheticElement extends Element {
    public double lifetime;
}
```

Like the discriminated subtypes of Ada, this approach constrains each object to a single variant at creation time, but this may not be a problem: while the class of a particular object never changes, class-type variables are references in Java and C#. A variable of type Element can easily refer to an object of class NaturalElement or SyntheticElement at run time. ■

CHECK YOUR UNDERSTANDING

39. What are *anonymous* unions and structs? What purpose do they serve? How is this related to the integration of variants with records in Pascal and its descendants?
40. What is a *tag* (*discriminant*) in a variant record? In a language like Ada or OCaml, how does it differ from an ordinary field?
41. Discuss the type safety problems that arise with variant records. How can these problems be addressed?
42. Summarize the rules that prevent access to inappropriate fields of variant records in OCaml and Ada.
43. Why might one wish to *constrain* a variable, so that it can hold only one variant of a type?
44. Explain how classes and inheritance can be used to obtain the effect of constrained variant records.

8 Composite Types

8.5.2 Dangling References

Memory access errors—dangling references, memory leaks, out-of-bounds access to arrays—are among the most common program bugs, and among the most difficult to find. Testing and debugging techniques for memory errors vary in when they are performed, how much they cost, and how conservative they are. Several commercial and open-source tools employ binary instrumentation (Section 16.2.3) to track the allocation status of every block in memory and to check every load or store to make sure it refers to an allocated block. These tools have proved to be highly effective, but they can slow a program several-fold, and may generate *false positives*—indications of error in programs that, while arguably poorly written, are technically correct. Many compilers can also be instructed to generate dynamic semantic checks for certain kinds of memory errors. Such checks must generally be fast (much less than 2× slowdown), and must never generate false positives. In this section we consider two candidate implementations of checks for dangling references.

Tombstones

EXAMPLE 8.68
Dangling reference detection with tombstones

Tombstones [Lom75, Lom85] allow a language implementation to catch all dangling references, to objects in both the stack and the heap. The idea is simple: rather than have a pointer refer to an object directly, we introduce an extra level of indirection (Figure C-8.17). When an object is allocated in the heap (or when a pointer is created to an object in the stack), the language run-time system allocates a tombstone. The pointer contains the address of the tombstone; the tombstone contains the address of the object. When the object is reclaimed, the tombstone is modified to contain a value (typically zero) that cannot be a valid address. To avoid special cases in the generated code, tombstones are also created for pointers to static objects. ■

For heap objects, it is easy to invalidate a tombstone when the program calls the deallocation operation. For stack objects, the language implementation must

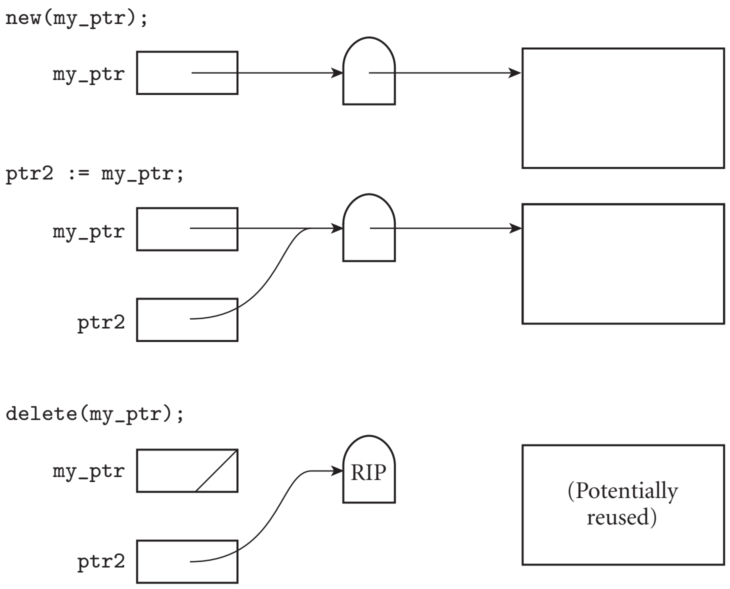


Figure 8.17 Tombstones. A valid pointer refers to a tombstone that in turn refers to an object. A dangling reference refers to an "expired" tombstone.

be able to find all tombstones associated with objects in the current stack frame when returning from a subroutine. One possible solution is to link all stack-object tombstones together in a list, sorted by the address of the stack frame in which the object lies. When a pointer is created to a local object, the tombstone can simply be added to the beginning of the list. When a pointer is created to a parameter, the run-time system must scan down the list and insert in the middle, to keep it sorted. When a subroutine returns, the epilogue portion of the calling sequence invalidates the tombstones at the head of the list, and removes them from the list.

Tombstones may be allocated from the heap itself or, more commonly, from a separate pool. The latter option avoids fragmentation problems, and makes allocation relatively fast, since the first tombstone on the free list is always the right size.

Tombstones can be expensive, both in time and in space. The time overhead includes (1) creation of tombstones when allocating heap objects or using a "pointer to" operator, (2) checking for validity on every access, and (3) double-indirection. Fortunately, checking for validity can be made essentially free on most machines by arranging for the address in an "invalid" tombstone to lie outside the program's address space. Any attempt to use such an address will result in a hardware interrupt, which the operating system can reflect up into the language run-time system. We can also use our invalid address, in the pointer itself, to represent the constant `nil`. If the compiler arranges to set every pointer to `nil` at elaboration time, then the hardware will catch any use of an uninitialized pointer. (This technique works without tombstones, as well.)

The space overhead for tombstones can be significant. The simplest approach is never to reclaim them. Since a tombstone is usually significantly smaller than the object to which it refers, a program will waste less space by leaving a tombstone around forever than it would waste by never reclaiming the associated object. Even so, any long-running program that continually creates and reclaims objects will eventually run out of space for tombstones. A potential solution, which we consider in Section 8.5.3, is to augment every tombstone with a *reference count*, and reclaim tombstones themselves when the reference count goes to zero.

Tombstones have a valuable side effect. Because of double-indirection, it is easy to change the location of an object in the heap. The run-time system need not locate every pointer that refers to the object; all that is required is to change the address in the tombstone. The principal reason to change heap locations is for *storage compaction*, in which all dynamically allocated blocks are "scooted together" at one end of the heap in order to eliminate external fragmentation. Tombstones are not widely used in language implementations, but the Macintosh operating system (versions 9 and below) used them internally, for references to system objects such as file and window descriptors. They also closely resemble the implementation used for *smart pointers* in the C++ standard library.

Locks and Keys

Locks and *keys* [FL80] are an alternative to tombstones. Their disadvantages are that they work only for objects in the heap, and they provide only probabilistic protection from dangling pointers. Their advantage is that they avoid the need to keep tombstones around forever (or to figure out when to reclaim them). Again the idea is simple: Every pointer is a tuple consisting of an address and a key. Every object in the heap begins with a lock. A pointer to an object in the heap is valid only if the key in the pointer matches the lock in the object (Figure C-8.18). When the run-time system allocates a new heap object, it generates a new key value. These can be as simple as serial numbers, but should avoid "common" values such as zero and one. When an object is reclaimed, its lock is changed to some arbitrary value (e.g., zero) so that the keys in any remaining pointers will not match. If the block is subsequently reused for another purpose, we expect it to be very unlikely that the location that used to contain the lock will be restored to its former value by coincidence. ■

EXAMPLE 8.69
Dangling reference detection with locks and keys

Like tombstones, locks and keys incur significant overhead. They add an extra word of storage to every pointer and to every block in the heap. They increase the cost of copying one pointer into another. Most significantly, they incur the cost of comparing locks and keys on every access (or every access that cannot be proven to be redundant). It is unclear whether the lock and key check is cheaper or more expensive than the tombstone check. A tombstone check may result in two cache misses (one for the tombstone and one for the object); a lock and key check is unlikely to cause more than one. On the other hand, the lock and key check requires a significantly longer instruction sequence on most machines.

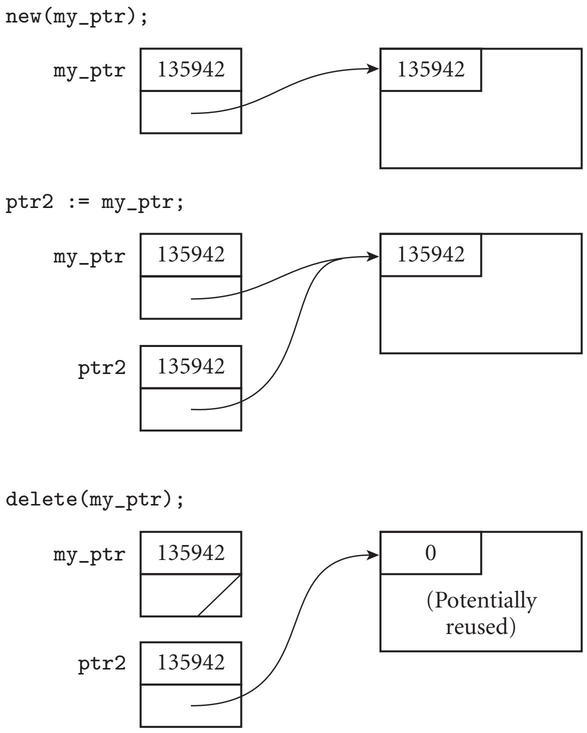


Figure 8.18 **Locks and keys.** A valid pointer contains a key that matches the lock on an object in the heap. A dangling reference is unlikely to match.

To minimize time and space overhead, most compilers do not by default generate code to check for dangling references. Most Pascal compilers allow the programmer to request dynamic checks, which are usually implemented with locks and keys. In most implementations of C, even optional checks are unavailable.

✓ CHECK YOUR UNDERSTANDING

45. What are *tombstones*? What changes do they require in the code to allocate and deallocate memory, and to assign and dereference pointers?
46. Explain how tombstones can be used to support *compaction*.
47. What are *locks* and *keys*? What changes do they require in the code to allocate and deallocate memory, and to assign and dereference pointers?
48. Explain why the protection afforded by locks and keys is only probabilistic.
49. Discuss the comparative advantages of tombstones and locks and keys as a means of catching dangling references.

8 Composite Types

8.7 Files and Input/Output

The first two subsections below are devoted to interactive and file-based I/O, respectively. Section C-8.7.3 then considers the common special case of text files.

8.7.1 Interactive I/O

On a modern machine, interactive I/O usually occurs through a graphical user interface (GUI: "gooey") system, with a mouse, a keyboard, and a bit-mapped screen that in turn support windows, menus, scrollbars, buttons, sliders, and so on. GUI characteristics vary significantly among, say, Microsoft Windows, the Macintosh, and Unix's X11; the differences are one of the principal reasons it is difficult to port applications across platforms.

Within a single platform, the facilities of a GUI system usually take the form of library routines (to create or resize a window, print text, draw a polygon, and so on). Input events (mouse move, button push, keystroke) may be placed in a queue that is accessible to the program, or tied to *event handler* subroutines that are called by the run-time system when the event occurs. Because the handler is triggered from outside, its activities must generally be *synchronized* with those of the main program, to make sure that both parties see a consistent view of any data shared between them. We will discuss events further in Section 9.6, and synchronization in Section 13.3.

A few programming languages—notably Smalltalk and Java—attempt to incorporate a standard set of GUI mechanisms into the language. The Smalltalk design team was part of the original group at Xerox's Palo Alto Research Center (PARC) that invented mouse-and-window based interfaces in the early 1970s. Unfortunately, while the Smalltalk GUI is successful within the confines of the language, it tends not to integrate well with the "look and feel" of the host system on which it runs. In a similar vein, Java's original GUI facilities (the Abstract

Window Toolkit—AWT) had something of a "least common denominator" look to them. Smalltalk's GUI is a fundamental part of the language; Java's takes the form of a standard set of library routines. The Java routines and their interface have evolved significantly over time; the more recent Swing and JavaFX libraries have "pluggable" look and feel, allowing them to integrate more easily with (and port more easily among) a variety of window systems.

The "parallel execution" of the program and the human user that characterizes interactive systems is difficult to capture in a functional programming model. A functional program that operates in a "batch" mode (taking its input from a file and writing its output to a file) can be modeled as a function from input to output. A program that interacts with the user, however, requires a very concrete notion of program ordering, because later input may depend on earlier output. If both input and output take the form of an ordered sequence of tokens, then interactive I/O can be modeled using lazy data structures, a subject we considered in Section 6.6.2. More general solutions can be based on the notion of *monads*, which use a functional notion of sequencing to model side effects. We will consider these issues again in Sections 11.5 and 11.8.

8.7.2 File-Based I/O

Persistent files are the principal mechanism by which programs that run at different times communicate with each other. A few language proposals (e.g., Argus [LS83] and χ [SH92]) allow ordinary variables to persist from one invocation of a program to the next, and a few experimental operating systems (e.g., Opal [CLFL94] and Hemlock [GSB$^+$93]) provide persistence for variables outside the language proper. In addition, some language-specific programming environments, such as those for Smalltalk and Common Lisp, provide a notion of *workspace* that includes persistent named variables. With coming advances in nonvolatile memory technology, such features may find their way into a larger number of languages. Historically, they have been more the exception than the rule. For the most part, data that need to outlive a particular program invocation have needed to reside in files.

Like interactive I/O, files can be incorporated directly into the language, or provided via library routines. In the latter case, it is still a good idea for the language designers to suggest a standard library interface, to promote portability of programs across platforms. The lack of such a standard in Algol 60 is widely credited with impeding the language's widespread use. One of the principal reasons to incorporate I/O into the language proper is to make use of special syntax. In particular, several languages, notably Fortran and Pascal, provide built-in I/O facilities in order to obtain type-safe "subroutines" that take a variable number of parameters, some of which may be optional.

Depending on the needs of the programmer and the capabilities of the host operating system, data in files may be represented in binary form, much as it is in memory, or as *text*. In a binary file, the number 1066_{10} would be represented by the 32-bit value 10000101010_2. In a text file, it would probably be represented by the character string `"1066"`. Temporary files are usually kept in binary form for the sake of speed and convenience. Persistent files are commonly kept in both forms. Text files are more easily ported across systems: issues of word size, byte order, alignment, floating-point format, and so on do not arise. Text files also have the advantage of human readability: they can be manipulated by text editors and related tools. Unfortunately, text files tend to be large, particularly when used to hold numeric data. A double-precision floating-point number occupies only eight bytes in binary form, but can require as many as 24 characters in decimal notation (not counting any surrounding white space). Text files also incur the cost of binary to text conversion on output, and text to binary conversion on input. The size problem can be addressed, at least for archival storage, by using data compression. Mechanisms to control text/binary conversion tend to be the most complicated part of I/O; we discuss them in the following subsection.

EXAMPLE 8.70
Files as a built-in type

When I/O is built into a language, files are usually declared using a built-in type constructor, as for example in Pascal:

```
var my_file : file of foo;
```

If I/O is provided by library routines, the library usually provides an opaque type to represent a file. In either case, each file variable is generally bound to an external, operating system–supported file by means of an *open* operation. In C, for example, one says

EXAMPLE 8.71
The `open` operation

```
my_file = fopen(path_name, mode);
```

The first argument to `fopen` is a character string that names the file, using the naming conventions of the host operating system. The second argument is a string that indicates whether the file should be readable, writable, or both, whether it should be created if it does not yet exist, and whether it should be overwritten or appended to if it does exist.

EXAMPLE 8.72
The `close` operation

When a program is done with a file, it can break the association between the file variable and the external object by using a *close* operation:

```
fclose(my_file);
```

In response to a call to `close`, the operating system may perform certain “finalizing” operations, such as unlocking an exclusive file (so that it may be used by other programs), rewinding a tape drive, or forcing the contents of buffers out to disk.

Most files, both binary and text, are stored as a linear sequence of characters, words, or records. Every open file then has a notion of *current position*: an implicit reference to some element of the sequence. Each `read` or `write` operation implicitly advances this reference by one position, so that successive operations access successive elements, automatically. In a *sequential* file, this automatic advance is the only way to change the current position. Sequential files usually correspond to media like printers and tapes, in which the current position has a physical representation (how many pages we've printed; how much tape is on each spool) that is difficult to change.

In other, *random-access* files, the programmer can change the current position to an arbitrary value by issuing a *seek* operation. In a few programming languages (e.g., Cobol and PL/I), random-access files (also called *direct* files) have no notion of current position. Rather, they are *indexed* on some key, and every `read` or `write` operation must specify a key. A file that can be accessed both sequentially *and* by key is said to be *indexed sequential*.

Random-access files usually correspond to media like solid-state flash drives or magnetic or optical disks, in which the current position can be changed with relative ease. Where tape drives (still widely used for archival storage) can take more than a minute to seek to a given position, modern disks take anywhere from 5 to 200 ms, depending on technology. (Note that 5 ms is still a very long time—10 million cycles on a 2 GHz processor.) Seeking on a solid-state device is essentially instantaneous. A few languages—notably Pascal—provide no random-access files, though individual implementations may support random access as a nonstandard language extension.

8.7.3 Text I/O

It is conventional to think of text files as consisting of a sequence of *lines*, each of which in turn consists of characters. In older systems, particularly those designed around the metaphor of punch cards, lines are reflected in the organization of the file itself. A `seek` operation, for example, may take a line number as argument. More commonly, a text file is simply a sequence of characters. Within this sequence, control (nonprinting) characters indicate the boundaries between lines. Unfortunately, end-of-line conventions are not standardized. In Unix and in modern versions of the Mac OS, each line of a text file ends with a *newline* ("control-J") character, ASCII value 10. (On "classic" Macs, each line ended with a *carriage return* ("control-M") character, ASCII value 13.) On Windows machines, each line ends with a carriage return/newline pair. Text files are usually sequential.

Despite the muddied conventions for line breaks, text files are much more portable and readable than binary files.[1] Because they do not mirror the structure of internal data, text files require extensive conversions on input and output. Issues to be considered include the base for integer values (and the representation of nondecimal bases); the representation of floating-point values (number

1 We are speaking here, of course, of plain-text ASCII or Unicode files. So-called "rich text" files, consisting of formatted text in particular fonts, sizes, and colors, perhaps with embedded graphics, are another matter entirely. Word processors typically represent rich text with a combination of binary and ASCII data, though ASCII-only standards such as Postscript, textual PDF, RTF, and XML can be used to enhance portability.

of digits, placement of decimal point, notation for exponent); the representation of enumerations and other nonnumeric, nonstring types; and positioning, if any, within columns (right and left justification, zero or white-space fill, "floating" dollar signs in Cobol). Some of these issues (e.g., the number of digits in a floating-point number) are influenced by the hardware, but most are dictated by the needs of the application and the preferences of the programmer.

In most languages the programmer can take complete control of input and output formatting by writing it all explicitly, using language or library mechanisms to read and write individual characters only. I/O at such a low level is tedious, however, and most languages also provide more high-level operations. These operations vary significantly in syntax and in the degree to which they allow the programmer to specify I/O formats. We illustrate the breadth of possibilities with examples from four imperative languages: Fortran, Ada, C, and C++.

Text I/O in Fortran

EXAMPLE 8.73
Formatted output in Fortran

In Fortran, we could write a character string, an integer, and an array of 10 floating-point numbers as follows:

```
character s*20
integer n
real r (10)
...
write (4, '(A20, I10, 10F8.2)'), s, n, r
```

In the `write` statement, the 4 indicates a *unit number*, which identifies a particular output file. The quoted, parenthesized expression is called a *format*; it specifies how the printed variables are to be represented. In this case, we have requested a 20-column ASCII string, a 10-column integer, and 10 eight-column floating-point numbers (with two columns of each reserved for the fractional part of the value). Fortran provides an extremely rich set of these *edit descriptors* for use inside of formats. Cobol, PL/I, and Perl provide comparable facilities, though with a very different syntax. ■

EXAMPLE 8.74
Labeled formats

Fortran allows a format to be specified indirectly, so it may be used in more than one input or output statement:

```
write (4, 100), s, n, r             ! 100 is the line number
...                                 ! of the format statement
100 format (A20, I10, 10F8.2)
```

It also allows formats to be created at run time, and stored in strings:

```
character(len=20) :: fmt
...
fmt = "(A20, I10, 10F8.2)"
...
write (4, fmt), s, n, r
```

If the programmer does not know, or does not care about, the precise allocation of columns to fields, the format can be omitted:

```
write (4, *), s, n, r
```

In this case, the run-time system will use default format conventions. ■

EXAMPLE 8.75
Printing to standard output

To write to the standard output stream (i.e., the terminal or its surrogate), the programmer can use the `print` statement, which resembles a `write` without a unit number:

```
print*, s, n, r          ! * means default format
```

For input, `read` is used both for standard input and for specific files; in the former case, the unit number is omitted, together with the extra set of parentheses:

```
read 100, s, n, r
...
read*, s, n, r           ! * means default format
```

The star may be omitted in Fortran 90. ■

In the full form of `read`, `write`, and `print`, additional arguments may be provided in the parenthesized list with the unit number and format. These can be used to specify a variety of additional information, including a label to which to jump on end-of-file, a label to which to jump on other errors, a variable into which to place status codes returned by the operating system, a set of labels (a "namelist") to attach to the output values, and a control code to override the usual automatic advance to the next line of the file. Because there are so many of these optional arguments, most of which are usually omitted, they are usually specified using *named* (keyword) parameter notation, a notion we defer to Section 9.3.3.

The variety of shorthand versions of `read`, `write`, and `print`, together with the fact that they operate on a variable number of program variables, makes it very difficult to cast them as "ordinary" subroutines. Fortran 90 provides optional and named parameters, but Fortran 77 does not, and even in Fortran 90 there is no way to define a subroutine with an *arbitrary* number of parameters.

Text I/O in Ada

Ada provides a suite of five standard library packages for I/O. The `sequential_IO` and `direct_IO` packages are for binary files. They provide generic file types that can be instantiated for any desired element type. The `IO_exceptions` and `low_level_IO` packages handle error conditions and device control, respectively. The `text_IO` package provides formatted input and output on sequential files of characters.

EXAMPLE 8.76
Formatted output in Ada

Using `text_IO`, our original three-variable Fortran output statement would look something like this in Ada:

```
s : array (1..20) of character;
n : integer;
r : array (1..10) of real;
...
set_output(my_file);
put(n, 10);
put(s);
for i in 1..10 loop put(r(i), 5, 2); end loop;
new_line;
```

In the `put` of an element of `r` (within the loop), the second parameter specifies the number of digits before the decimal point, rather than the width of the entire number (including the decimal point), as it did in Fortran. The `put` of `s` will use the string's natural length. If a different length is desired, the programmer will have to write blanks or `put` a substring explicitly. If precise output positioning is not desired for the integers and real numbers, the extra parameters in their `put` calls can be omitted; in this case the run-time system will use standard defaults. The programmer can use additional library routines to change these defaults if desired. A call to `set_output` invokes a similar mechanism: it changes the default notion of output file. ■

EXAMPLE 8.77
Overloaded `put` routines

There are two overloaded forms of `put` for every built-in type. One takes a file name as its first argument; the other does not. The last five lines above could have been written

```
put(my_file, n, 10);
put(my_file, s);
for i in 1..10 loop put(my_file, r(i), 5, 2); end loop;
new_line(my_file);
```

The programmer can of course define additional forms of `get` and `put` for arbitrary user-defined types. All of these facilities rely on standard Ada mechanisms; in contrast to Fortran, no support for I/O is built into the language itself. ■

Text I/O in C

C provides I/O through a library package called `stdio`; as in Ada, no support for I/O is built into the language itself. Many C implementations, however, build knowledge of I/O functions into the compiler, so it can issue warnings when arguments appear to be used incorrectly.

EXAMPLE 8.78
Formatted output in C

Our example output statement would look something like this in C:

```
char s[20];
int n;
double r[10];
...
fprintf(my_file, "%20s%10d", s, n);
for (i = 0; i < 10; i++) fprintf(my_file, "%8.2f", r[i]);
fprintf(my_file, "\n");
```

The arguments to `fprintf` are a file, a format string, and a sequence of expressions. The format string has capabilities similar to the formats of Fortran, though the syntax is very different. In general, a format string consists of a sequence of characters with embedded "placeholders," each of which begins with a percent sign. The placeholder `%20s` indicates a 20-character string; `%d` indicates an integer in decimal notation; `%8.2f` indicates an 8-character floating-point number, with two digits to the right of the decimal point. ■

EXAMPLE 8.79
Text in format strings

As in Fortran, formats can be computed and stored in strings, and a single `fprintf` statement can print an arbitrary number of expressions. As in Ada, an explicit `for` loop is needed to print an array. Commonly the format string also contains labeling text and white space:

```
strcpy(s, "four");                    /* copy "four" into s */
n = 20;
char *fmt = "%s score and %d years ago\n";
fprintf(my_file, fmt, s, n);
```

A percent sign can be printed by doubling it:

```
fprintf(my_file, "%d%%\n", 25);     /* prints "25%" */
```
■

EXAMPLE 8.80
Formatted input in C

Input in C takes a similar form. The `fscanf` routine takes as argument a file, a format string, and a sequence of pointers to variables. In the common case, every argument after the format is a variable name preceded by a "pointer to" operator:

```
fscanf(my_file, "%s %d %lf", &s, &n, &r[0]);
```

In this call, the `%s` placeholder will match a string of maximal length that does not include white space. If this string is longer than 20 characters (the length of `s`), then `fscanf` will write beyond the end of the storage for the string. (This weakness in `scanf` is one of the sources of the so-called "buffer overflow" bugs discussed in Sidebar 8.7. It can be avoided in this example by replacing the `%s` specifier with `%19s`, which will cause `fscanf` to move at most 19 bytes, plus a terminating `NUL`.) The three-character `%lf` placeholder informs the library routine that the corresponding argument is a `double`; the 2-character sequence `%f` would read into a `float`.[2] Accidentally using a placeholder for the wrong size variable is a common error in older implementations of C; forgetting the ampersand on a trailing argument is another. While such mistakes will often be caught

2 C's `doubles` are double-precision IEEE floating-point numbers in most implementations; `floats` are usually single precision. The lack of safety for `%s` arguments is only one of several problems with `fscanf`. Others include the inability to "skip over" erroneous input, and undefined behavior when there is insufficient input. Instead of `fscanf`, seasoned C programmers tend to use `fgets`, which reads (length-limited) input into a string, followed by manual parsing using `strtol` (string-to-`long`), `strtod` (string-to-`double`), and so on.

by a modern C compiler with special-case knowledge of `fscanf`, they would always be caught in a language with type-safe I/O. Note that we have read a single element of `r`; as with `fprintf`, a `for` loop would be needed to read the whole array. ■

We have noted above that the I/O routines of Fortran and Pascal are built into the language largely to permit them to take a variable number of arguments. We have also noted that moving I/O into a library in Ada forces us to make a separate call to `put` for every output expression. So how do `fprintf` and `fscanf` work? It turns out that C permits functions with a variable number of parameters (we will discuss such functions in more detail in Section 9.3.3). Unfortunately, the types of trailing parameters are unspecified, which makes compile-time type checking of variable-length argument lists impossible in the general case. Moreover, the lack of run-time type descriptors in C precludes run-time checking as well. At the same time, because the C library (including `fprintf` and `fscanf`) is part of the language standard, special knowledge of these routines can be built into the compiler—and often is: while the I/O routines of C are formally defined as "ordinary" functions, they are typically implemented in the same way as their analogues in Fortran and Pascal. As a result, C compilers will often provide good error diagnostics when the arguments to `fprintf` or `fscanf` do not match the format string.

To simplify I/O to and from the standard input and output streams, `stdio` provides routines called `printf` and `scanf` that omit the initial arguments of `fprintf` and `fscanf`. To facilitate the formatting of strings *within* a program, `stdio` also provides routines called `sprintf` and `sscanf`, which replace the initial arguments of `fprintf` and `fscanf` with a pointer to an array of characters. The `sscanf` function "reads" from this array; `sprintf` "writes" to it. Fortran 90 provides similar support for intraprogram formatting through so-called *internal files*.

Text I/O in C++

As a descendant of C, C++ supports the `stdio` library described in the previous subsection. It also supports a new I/O library called `iostream` that exploits the object-oriented features of the language. The `iostream` library is more flexible than `stdio`, provides arguably more elegant syntax (though this is a matter of taste), and is completely type safe.

C++ *streams* use operator overloading to co-opt the << and >> symbols normally used for bit-wise shifts. The `iostream` library provides an overloaded version of << and >> for each built-in type, and programmers can define versions for new types. To print a character string in C++, one writes

EXAMPLE 8.81
Formatted output in C++

```
my_stream << s;
```

To output a string and an integer one can write

```
my_stream << s << n;
```

This code requires that `my_stream` be an instance of the `ostream` (output stream) class defined in the `iostream` library. The << operator is syntactic sugar for the "operator function" `ostream::operator<<`, as described in Section 3.5.2. Because << associates left-to-right, the statement above is equivalent to

```
(my_stream.operator<<(s)).operator<<(n);
```

The code works because `ostream::operator<<` returns a reference to its first argument as its result (as we shall see in Section 9.3.1, C++ supports both a value model and a reference model for variables). ■

EXAMPLE 8.82
Stream manipulators

As shown so far, output to an `ostream` uses default formatting conventions. To change conventions, one may embed so-called *stream manipulators* in a sequence of << operations. To print `n` in octal notation (rather than the default decimal), we could write

```
my_stream << oct << n;
```

To control the number of columns occupied by `s` and `n`, we could write

```
my_stream << setw(20) << s << setw(10) << n;
```

The `oct` manipulator causes the stream to print all subsequent numeric output in octal. The `setw` manipulator causes it to print its next string or numeric output in a field of a specified minimum width (behavior reverts to the default after a single output). ■

The `oct` manipulator is declared as a function that takes an `ostream` as a parameter and produces a reference to an `ostream` as its result. Because it is not followed by empty parentheses, the occurrence of `oct` in the output sequence above is *not* a call to `oct`; rather, a reference to `oct` is passed to an overloaded version of << that expects a manipulator function as its right-hand argument. This version of << then calls the function, passing the stream (the left-hand argument of <<) as argument.

The `setw` manipulator is even trickier. It is declared as a function that returns a reference to what we might call an "object closure"—an object containing a reference to a function and a set of arguments. In this particular case, `setw(20)` is a call to a *constructor* function that returns a closure containing the number 20 and a pointer to the `setw` manipulator. (We will discuss constructors in detail in Section 10.3, and object closures in Section 3.6.3.) The `iostream` library provides an overloaded version of << that expects an object closure as its right-hand argument. This version of << calls the function inside the closure, passing it as arguments the stream (the left-hand argument of <<) and the integer inside the closure.

EXAMPLE 8.83
Array output in C++

The `iostream` library provides a wealth of manipulators to change the formatting behavior of an `ostream`. Because C++ inherits C's handling of pointers and arrays, however, there is no way for an `ostream` to know the length of an array. As a result, our full output example still requires a `for` loop to print the `r` array:

```
char s[20];
int n;
double r[10];
...
my_stream << setw(20) << s << setw(10) << n;
for (i = 0; i < 10; i++)
    my_stream << setiosflags(ios::fixed)
        << setw(8) << setprecision(2) << r[i];
my_stream << "\n";
```

Here the manipulators in the output sequence in the `for` loop specify fixed format (rather than scientific) for floating-point numbers, with a field width of eight, and two digits past the decimal point. The `setiosflags` and `setprecision` manipulators change the default format of the stream; the changes apply to all subsequent output. ■

EXAMPLE 8.84
Changing default format

To avoid calling stream manipulators repeatedly, we could modify our example as follows:

```
my_stream.flags(my_stream.flags() | ios::fixed);
my_stream.precision(2);
for (i = 0; i < 10; i++) my_stream << setw(8) << r[i];
```

To facilitate the restoration of defaults, the `flags` and `precision` functions return the previous value:

```
ios::fmtflags old_flags = my_stream.flags(my_stream.flags() | ios::fixed);
int old_precision = my_stream.precision(2);
for (i = 0; i < 10; i++) my_stream << setw(8) << r[i];
my_stream.flags(old_flags);
my_stream.precision(old_precision);
```
■

Formatted input in C++ is analogous to formatted output. It uses `istreams` instead of `ostreams`, and the `>>` operator instead of `<<`. It also supports a suite of manipulators comparable to those for output. I/O on the standard input and output streams does not require different functions; the programmer simply begins an input or output sequence with the standard stream name `cin` or `cout`. (In keeping with C tradition, there is also a standard stream `cerr` for error messages.) To support intraprogram formatting of character strings, the `strstream` library provides `istrstream` and `ostrstream` object classes that are derived from `istream` and `ostream`, and that allow a stream variable to be bound to a string instead of to a file.

✓ **CHECK YOUR UNDERSTANDING**

50. Explain the differences between interactive and file-based I/O, between temporary and persistent files, and between binary and text files. (Some of this information is in the main text.)
51. What are the comparative advantages of *text* and *binary* files?
52. Describe the end-of-line conventions of Unix, Windows, and Macintosh files.
53. What are the advantages and disadvantages of building I/O into a programming language, as opposed to providing it through library routines?
54. Summarize the different approaches to text I/O adopted by Fortran, Ada, C, and C++.
55. Describe some of the weaknesses of C's `scanf` mechanism.
56. What are *stream manipulators*? How are they used in C++?

8 Composite Types

8.9 Exercises

8.23 In Example 6.70 we described a programming idiom in which an iterator takes a “loop body” function as argument, and applies it to every element of a given container or set. Show how to use this idiom in ML to apply a function to every element of the tree in Example 11.39. Write versions of your iterator for preorder, inorder, and postorder traversals.

8.24 Show how unions can be used in C to interpret the bits of a value of one type as if they represented a value of some other type. Explain why the same technique does not work in Ada. After consulting an Ada manual, describe how an `unchecked` pragma can be used to get around the Ada rules.

8.25 Are variant records a form of polymorphism? Why or why not?

8.26 Learn the details of variant records in Pascal.

(a) You may have noticed that the language does not allow you to pass the tag field of a variant record to a subroutine by reference. Why not?

(b) Explain how you might implement dynamic semantic checks to catch references to uninitialized fields of a tagged variant record. Changing the value of the tag field should cause all fields of the variant part of the record to become uninitialized. Suppose you want to avoid adding flag fields within the record itself (e.g., to avoid changing the offsets of fields in a systems program). How much harder is your task?

(c) Explain how you might implement dynamic semantic checks to catch references to uninitialized fields of an *untagged* variant record. Any assignment to a field of a variant should cause all fields of other variants to become uninitialized. Any assignment that changes the record from one variant to another should also cause all other fields of the new variant to be uninitialized. Again, suppose you want to avoid adding flag fields within the untagged record itself. How much harder is your task?

8.27 We noted in Section C-8.1.3 that Ada requires the variant portions of a record to occur at the end, to save space when a particular record is constrained to have a comparatively small variant part. Could a compiler rearrange fields to achieve the same effect, without the restriction on the declaration order of fields? Why or why not?

8.28 In Example 8.52 we noted that reference counts can be used to reclaim tombstones, failing only when the programmer neglects to manually delete the object to which a tombstone refers. Explain how to leverage this observation to catch memory leaks at run time. Does your solution work in all cases? Explain.

8.29 In Section 8.5.3 we introduced the notion of *smart pointers* in C++. Learn how these are implemented, and write an explanation. Discuss the relationship to tombstones.

8.30 Rewrite Example C-8.80 using `fgets`, `strtol`, `strtod`, etc. (read the `man` pages), so that it is guaranteed not to result in buffer overflow.

8.31 The output routines of several languages (e.g., `println` in Swift) give special treatment to ends of lines. By contrast, C's `printf` does not; it treats newlines and carriage returns the same as any other character. What are the comparative advantages of these approaches? Which do you prefer? Why?

8 Composite Types

8.10 Explorations

8.39 Research the history of smart pointers (Section 8.5.3) in C++, including the `unique_ptr`, `shared_ptr`, and `weak_ptr` of C++11; the `auto_ptr` of C++98, and the various pointer classes of the popular Boost library. How has the standard set of pointers evolved over time? What accounts for the changes? Do you consider the current mechanisms an adequate replacement for automatic garbage collection? Why or why not?

8.40 Find a Cobol manual and learn about the language's facilities for text I/O. Prepare a written comparison of those facilities to those of the languages described in Section C-8.7.3.

8.41 If you were designing the text I/O facilities for a new programming language, what approach would you take? In particular, do you believe that I/O should be a built-in part of the language, or should it be handled by library routines?

9 Subroutines and Control Abstraction

9.2.1 Displays

EXAMPLE 9.56
Nonlocal access using a display

As noted in the main text, a display is an embedding of the static chain into an array. The *j*th element of the display contains a reference to the frame of the most recently active subroutine at lexical nesting level *j*. The first element of the display is thus a reference to the frame of some subroutine *S* nested directly inside the main program; the second element is a reference to the frame of a routine that is nested inside of *S*, and so forth, until we reach the currently active routine. Figure C-9.7 contains an example. ■

If the display is stored in memory, then a nonlocal object can be loaded into a register with two memory accesses: one to load the display element into a register, the second to load the object. On a machine with a large number of registers, one might be tempted to reduce the overhead to only one memory access by keeping the entire display in registers, but that would probably be a bad idea: display elements tend to be accessed much less frequently than other things (e.g., local variables) that might be kept in the registers instead.

Maintaining the Display

Maintenance of a display is slightly more complicated than maintenance of a static chain, but not by much. Perhaps the most obvious approach would be to maintain the static chain as usual, and simply fill the display at procedure entry and exit, by walking down the chain. In most cases, however, the following (much faster) scheme suffices: when calling a subroutine at lexical nesting level *j*, the callee saves the current value of the *j*th display element into the stack, and then replaces that element with a copy of its own (newly created) frame pointer. Before returning, it restores the old element. Why does this mechanism work? As with static chains, there are two cases to consider:

1. The callee is nested (directly) inside the caller. In this case the caller and the callee share all display elements up to the current level. Putting the callee's frame pointer into the display simply extends the current level by one. It is conceivable that the old value needn't be saved, but in general there is no way

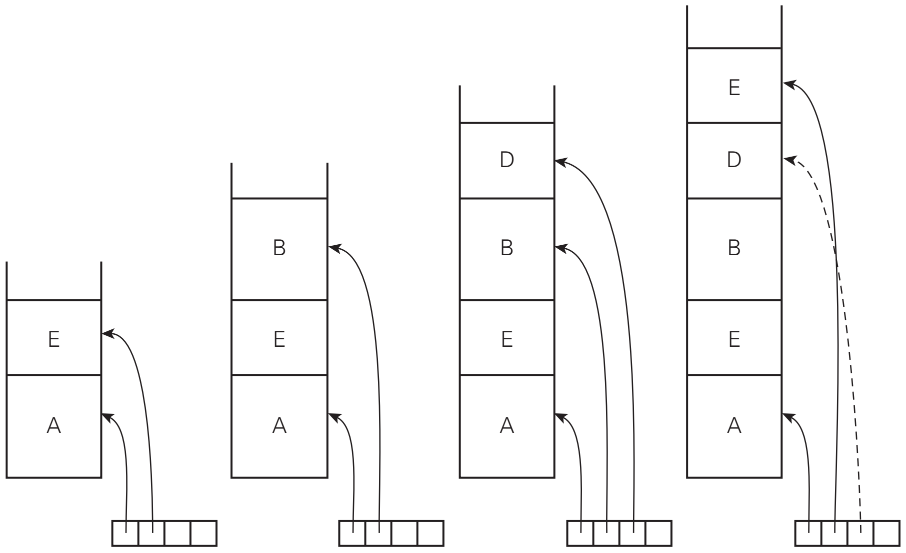


Figure 9.7 **Nonlocal access using a display.** The stack configurations, from left to right, illustrate the contents of the display (at bottom) for a sequence of subroutine calls, assuming the lexical nesting of Figure 9.1. Display elements beyond that of the currently executing subroutine are not used.

to tell. The caller itself might have been called by code that is very deeply nested, and that is counting on the integrity of a very deep display, in which case the old display element *will* be needed. A smart compiler may be able to avoid the save in certain circumstances.

2. The callee is at lexical nesting level j, $k \geq 0$ levels out from the caller. In this case the caller and callee share all display elements up through $j - 1$. The caller's

DESIGN & IMPLEMENTATION

9.8 Lexical nesting and displays

Because the display is a fixed-size array, compilers that use a display to implement access to nonlocal objects generally impose a limit (the size of the display) on the maximum depth to which subroutines may be nested. If this limit is larger than, say, five or six, it is unlikely that any programmer will ever wish for more. Note that the display does not eliminate the need for a frame pointer. Because local variables are accessed so often, it is important to have the address of the current frame in a register, where it can be used for displacement-mode addressing. Similarly, on a RISC processor, where a 32-bit address will not fit in one instruction, it is important to maintain a base register for the most commonly accessed global variables as well.

entry at level j is different from the callee's, so the callee must save it before storing its own frame pointer. If the callee in turn calls a routine at level $j + 1$, that routine will change another element of the display, but all old elements will be restored before they are needed again.

If the callee is a leaf routine then the display can be left intact; no one will use the element corresponding to the callee's nesting level before control returns to the caller.

Closures

A subroutine that is passed as a parameter, stored in a variable, or returned from a function must be called through some sort of *closure* (Section 3.6) that captures the referencing environment. In a language implementation based on static chains, a closure can be represented as a ⟨code address, static link⟩ pair. Displays are not as simple. A standard technique is to create two "entry points"—starting addresses—for every subroutine. One of these is for "normal" calls, the other for calls through closures. When a closure is created, it contains the address of the alternative entry point. The code at that entry point saves elements 1 through j of the display into the stack (it will have to create a larger-than-normal stack frame in order to do this), and then replaces those elements with values taken from (or calculated from) the closure. The alternative entry then makes a nested call to the main body of the subroutine (it skips the code immediately following the normal entry—the code that creates the normal stack frame and updates the display). When the subroutine returns, it comes back to the code of the alternative entry, which restores the old value of the display before returning to the actual caller.

More space-conserving implementations of display-based closures are possible (see Exercise C-9.26), but with higher run-time overhead.

Comparison to Static Chains

In general, maintaining a display is slightly more expensive than maintaining a static chain, though the comparison is not absolute. In the usual case, passing a static link to a called routine requires $k \geq 0$ load instructions in the caller, followed by one store instruction in the callee (to place the static link at the appropriate offset in the stack frame). The store may be skipped in leaf routines, assuming that a register is available to hold the link as long as it is needed. No overhead is required to maintain the static chain when returning from a subroutine. With a display, a nonleaf callee requires two loads and three stores (1 + 2 in the prologue and 1 + 1 epilogue) to save and restore display elements. Because the callee does all the work, displays may save a little bit on code size, compared to static chains. As noted above, displays significantly complicate the creation and use of closures.

The original advantage of displays—reduced cost for access to objects in outer scopes—seems less clear today than once it did. In fact, while displays were popular in the CISC compilers of the 1970s and 1980s, they are less common in recent compilers. Most programs don't nest subroutines more than two or three levels deep, so static chains are seldom very long, and variables in surrounding scopes tend not to be accessed very often. If they *are* accessed often, common subexpression optimizations (to be discussed in Chapter 17) are likely to ensure that a pointer to the appropriate frame remains in a register.

Some language designers have argued that the development of object-oriented programming (the subject of Chapter 10) has eliminated the need for nested subroutines [Han81]. Others might even say that the success of C has shown such routines to be unneeded. Without nested subroutines, of course, the choice between static chains and displays is moot.

✓ CHECK YOUR UNDERSTANDING

44. Describe how we access an object at lexical nesting level *k* in a language implementation based on displays.
45. Why isn't the display typically kept in registers?
46. Explain how to maintain the display during subroutine calls.
47. What special concerns arise when creating closures in a language implementation that uses displays?
48. Summarize the tradeoffs between displays and static chains. Describe a program for which displays will result in faster code. Describe another for which static chains will be faster.

9 Subroutines and Control Abstraction

9.2.2 Stack Case Studies: LLVM on ARM; `gcc` on x86

To make stack management a bit more concrete, we present a pair of case studies: Apple's LLVM-based C compiler for the iPhone (ARM) and the GNU compiler suite for 32- and 64-bit x86. Both examples follow the general scheme outlined in Section 3.2.2, with differences in details that reflect the history of the respective compilers and the architecture of the target machines.

LLVM on ARM

An overview of the ARM instruction set architecture (ISA) can be found in Section C-5.4.5. For the sake of interoperability, ARM Ltd. publishes a standard for subroutine calling sequences that allows code from different vendors and compilers to link and run together. The standard has several variants, reflecting hardware features (Thumb mode, floating-point or vector instructions and registers, dynamic linking) that may or may not be present on a given processor or in its software environment. We focus here on the conventions adopted by Apple's C compiler for the iPhone and iPad (version 4.2), at optimization level `-O2`. The Apple compiler uses the 32-bit ARM back end (version 3.2svn) of the LLVM compiler suite. Given the level of detail in ARM's standard, code produced by other compilers is likely to be quite similar. Note, however, that the conventions for 64-bit code are very different; they are not documented here.

As noted in Section C-5.4.5, register `r14` (also known as `lr`) is special-cased by the ISA to receive the return address in subroutine call (`bl`—branch-and-link) instructions. Register `r13` (also known as `sp`) is similarly reserved for use as the stack pointer. It is not modified by `bl` instructions, but several variants of `push` and `pop`, which do update `sp`, are commonly part of the subroutine calling sequence. Some compilers for ARM, though not all, dedicate a third register by convention for use as a frame pointer; LLVM uses `r7` for this purpose.

EXAMPLE 9.57
LLVM/ARM stack layout

A typical LLVM/ARM stack frame appears in Figure C-9.8. The `sp` register points to the *last used* location in the stack (note that some compilers aim the

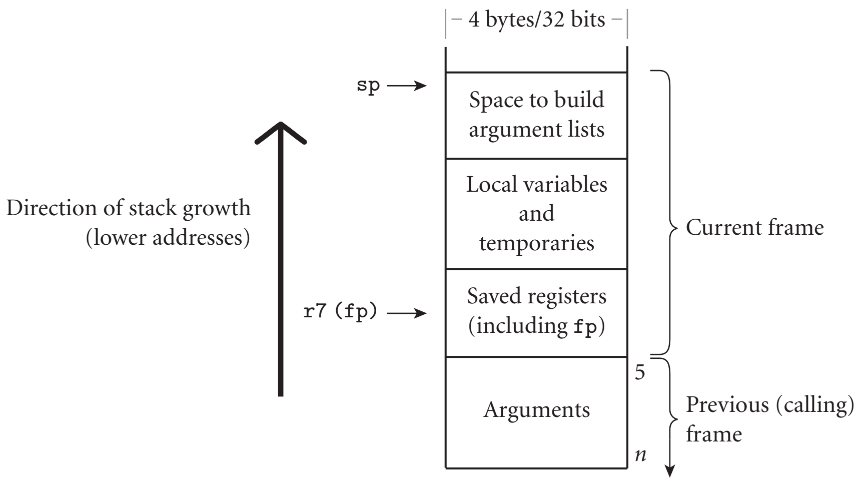

4 bytes/32 bits
sp
Space to build argument lists
Local variables and temporaries
Saved registers (including fp)
Arguments
r7 (fp)
Direction of stack growth (lower addresses)
Current frame
Previous (calling) frame
5
n

Figure 9.8 **Layout of the subroutine call stack for Apple's LLVM-based C compiler for ARM, running in 32-bit mode.** As in Figure 9.2, lower addresses are toward the top of the page.

pointer at the *first unused* location). ARM's subroutine calling standard requires that the stack always be word-aligned (`sp` mod 4 = 0). At an external call (to a subroutine in a different compilation unit) it must be double-word aligned (`sp` mod 8 = 0).

The first four arguments to a subroutine are always passed in registers. Additional arguments may be passed on the stack, with the last argument in the deepest location. Space for stack-based arguments is considered a part of the calling routine. If the current routine is not a leaf, space for any stack-based arguments it needs to pass to additional routines is preallocated, at the top (lowest-addressed-end) of the frame, as part of the subroutine prologue.

Space for local variables and for any temporary values that will not fit in registers is allocated in the middle of the frame. If the subroutine ever applies an address-of operator (`&` in C) to a low-numbered argument (one that will have been passed in a register), or if it passes such an argument to another routine by reference, the compiler creates a local variable to hold the argument, and initializes it with the value passed in the register.

Any callee-saves registers that may be overwritten by the current routine are saved at the bottom of the frame, directly beyond any stack-based arguments. The frame pointer (`r7`) is typically among these. LLVM arranges for the current `fp` to point to the location of the saved `fp`. ■

Argument Passing Conventions Arguments and locals of the current subroutine are accessed via offsets from the `fp`. Arguments in the process of being passed to the next routine are assembled at the top of the frame, and are accessed via offsets from the `sp`. The first four arguments are passed in registers `r0`–`r3`. A double-precision floating-point number is divided into two 32-bit halves, and passed as if it were two integers. (Some other ARM compilers pass floating-point arguments in the floating-point registers.) Records (`structs`) that appear early in the argument list may also be split into 32-bit pieces, and passed in multiple registers. An argument may be split between registers and the stack, if part but not all of it will fit in registers.

The argument build area at the top of the frame is designed to be large enough to hold the largest argument list that may be passed to any called routine. This convention may waste a bit of space in certain cases, but it ensures that arguments never need to be "pushed" in the usual sense of the word: the `sp` does not change when they are placed into the stack.

Return values up to 4 bytes in length occupy register `r0`. Double-word scalar return values occupy register pair `r0`–`r1`; quad-word scalar return values occupy registers `r0`–`r3`. Record-type return values of more than four bytes are placed in memory, at a location chosen by the caller and passed as an extra, hidden argument. If the return value is to be assigned immediately into a variable (e.g., `x = foo()`), the caller can simply pass the address of the variable. If the value is to be passed in turn to another subroutine, the caller can pass the appropriate address within its own argument build area. (Writing the return value into this space will probably destroy the returning function's own arguments, but that's fine in the absence of call-by-value/result: at this point the arguments are no longer needed.) Finally, though one doesn't see this idiom often (and most languages don't support it), C allows the caller to extract a field directly from the return value of a function (e.g., `x = foo().a + y;`); in this case the caller must pass the address

DESIGN & IMPLEMENTATION

9.9 Leveraging `pc = r15`

Because ARM assigns a register number to the program counter, that counter can be read and written (almost) like any other register. Writes to the `pc` cause a branch in control. This convention, together with the choice of `lr` = `r14` and `pc` = `r15`, enables an interesting optimization. If a subroutine is not a leaf (i.e., it calls another routine), `lr` will be among the registers saved at the bottom of the frame. If we suppose, for concreteness, that the subroutine plans to overwrite callee-saves registers `r4` and `r5`, and we know that we need to update the frame pointer (`r7`), then the subroutine prologue is likely to contain a `push {r4, r5, r7, lr}` instruction. This instruction stores the registers in sorted order, with the highest-numbered register (in this case, `lr`) at the highest address—deepest in the stack. One might naturally expect the epilogue to contain a symmetric `pop {r4, r5, r7, lr}` instruction, followed immediately by `bx lr` (branch to location in `lr`). But since the `pc` and `lr` have adjacent register numbers, the compiler can—and typically does—achieve the same result with a single `pop {r4, r5, r7, pc}` instruction.

of a temporary location within the "local variables and temporaries" part of its stack frame.

ARM and Thumb Mode Switching One of the more unusual features of the 32-bit ARM ISA (as described in Section C-5.4.5) is the presence of two separate instruction encodings. As on most RISC machines, the A32 encoding represents each instruction with 32 bits. The alternative T32 encoding, also known as "Thumb," represents the most common instructions in only 16 bits; the resulting improvement in code density can be important in embedded applications. While the two encodings are quite different (and in particular, T32 is not a subset of A32), program fragments that use different encodings can be linked into a single program.

To switch from one format to another, the program uses special `bx` (branch and exchange instruction set) and `blx` (branch with link and exchange instruction set) instructions. When the target address is statically known, the assumption is that the programmer knows that the source and target encodings are different, so the processor needs to change modes in the course of performing the branch. When the target address is in a register (as it will be when returning from a subroutine, or when calling through a pointer, a virtual method table, or a closure), ARM exploits the fact that instructions never appear at an odd address (T32 instructions are always word aligned; A32 instructions are always longword aligned). Because the least significant bit of the target address must always be 0, this bit can be used in the register to specify the target instruction set: 0 means A32; 1 means T32.

EXAMPLE 9.58
LLVM/ARM calling sequence

Calling Sequence Details The calling sequence to maintain the LLVM/ARM stack is as follows. The caller

1. saves (into the "local variables and temporaries" part of its frame) any caller-saves registers whose values are still needed
2. puts up to four small arguments (or "chunks" of larger arguments) into registers `r0`–`r3`
3. stores the remaining arguments into the argument build area at the top of the current frame
4. performs a `bl` or `blx` instruction, which puts the return address in register `lr`, jumps to the target address, and optionally changes instruction set encoding

On 32-bit ARM, the caller-saves registers are just the ones that are used for arguments—namely, `r0`–`r3`. In a language with nested subroutines (not supported by Apple's compiler), the caller would need to place the static link into another register immediately before performing the `bl` or `blx`.

In its prologue, the callee

1. pushes any necessary registers onto the stack
2. initializes the frame pointer by adding an appropriate small constant to the `sp`, placing the result in `r7`

3. subtracts enough from the sp to make space for local variables, temporaries, and the argument build area at the top of the stack, rounding down to a lower address if necessary to ensure that these objects have appropriate alignment

Saved registers include (a) the frame pointer, r7 (assuming the current routine needs a frame pointer of its own); (b) any callee-saves registers (r4–r6 and r8–r11) whose values may be changed before returning; and (c) the link register, lr, if the current routine is not a leaf, or if it uses lr as an additional temporary.

In its epilogue, immediately before returning, the callee

1. places the function return value (if any) into r0–r3 or memory, as appropriate
2. subtracts a small constant from r7, placing the result in sp; this effectively deallocates the bulk of the frame
3. pops saved registers from the stack, with the pc taking the place held by lr in the corresponding save in the prologue; this has the side effect of branching back to the caller (see Sidebar C-9.9)

Finally, if appropriate, the caller moves the return value to wherever it is needed. Caller-saves registers are restored lazily over time, as their values are needed.

To support the use of symbolic debuggers, the compiler generates a wealth of symbol table information, in the open-source DWARF format [DWA10]. It embeds this information into the object file. The information is most accurate when the program is compiled without any code improvement (-O0). For each subroutine, the information includes the starting and ending addresses of the routine; the name, type, and location (register name or frame pointer offset) of every formal parameter and local variable; the set of instructions corresponding to each line of source code; the size and layout of the stack frame; and a list of which registers were saved. ■

gcc on x86

To illustrate the differences among compilers and architectures, our second case study considers the GNU compiler collection (gcc, version 4.8.1) on the x86. We begin with 32-bit code and then explain the differences that obtain on 64-bit machines. Our example again focuses mostly on C, which acts as sort of a "lowest common denominator" among high-level languages. We also consider nested subroutines and closures, however, since these appear in some of the collection's supported languages.

An overview of the x86-32 ISA appears in Section C-5.4.5. Given the machine's CISC heritage and the comparatively small number of registers (only six are available for general-purpose use), all arguments are passed on the stack when running in 32-bit mode. To give the compiler the freedom to evaluate arguments out of order when desired, recent versions of gcc employ an argument build area similar to that of the LLVM case study. Unlike LLVM, recent versions of gcc omit the use of a separate frame pointer by default, making register ebp (rbp in 64-bit mode) available for other purposes; exceptions occur when specified by the programmer (using the `-no-omit-frame-pointer` command-line switch), when compiling at optimization levels `-O0` and `-O1`, when a subroutine has a local variable whose size is not known at compile time (Figure 8.7), or when a subroutine calls `alloca` (a legacy mechanism to create temporary space within the current stack frame). Historically, omission of the frame pointer made it difficult or even impossible for symbolic debuggers to perform a "backtrace" operation (identifying the frames of calling routines), but this limitation has been removed with modern debugging standards like DWARF.

Calling sequences for the x86 vary from vendor to vendor, and have evolved considerably over time, as changes in microarchitecture changed performance tradeoffs. Most modern sequences use the `call` and `ret` instructions. The former pushes the return address onto the stack, updating the `sp`, and branches to the called routine. The latter pops the return address off the stack, again updating the `sp`, and branches back to the caller. Several additional, more complex instructions, retained for backward compatibility, are typically not generated by modern compilers, because they were designed for calling sequences with an explicit display and without an argument build area, or because they don't pipeline as well as equivalent sequences of simpler instructions.

EXAMPLE 9.59
gcc/x86-32 stack layout

Argument Passing Conventions Figure C-9.10 shows a stack frame for the x86-32. As in the LLVM case study, the `sp` points to the last used location on the stack. Arguments in the process of being passed to another routine are accessed via offsets from the `sp`; everything else is accessed via offsets from the `fp`, if present—otherwise the `sp`. All arguments are passed in the stack. In languages (Ada, in particular) that permit nested subroutines, register `ecx` is used to pass the static link. If the current routine has at least one lexically nested child and is itself lexically nested in some parent, then a copy of the static link will be saved into the stack just above (at a lower address than) the area used for local variables and temporaries. When a nested routine is running, its own static link will point to the saved link in this current routine, or to the local variables and temporaries, if this current routine is outermost.

Functions return integer or pointer values in register `eax`. Floating-point values are returned in the first of the "x87" floating-point registers, `st(0)`. Composite values (records, arrays, etc.) of 8 bytes or less are returned in the register pair `eax–edx`, as are "`long long`" (64-bit) integers. For larger return values (records, arrays, etc.), the compiler passes a hidden first argument (on the stack) whose value is the address at which the return value should be written. ■

EXAMPLE 9.60
gcc/x86-32 calling sequence

Calling Sequence Details The calling sequence to maintain the gcc/x86-32 stack is as follows. The caller

1. saves (into the "local variables and temporaries" part of its frame) any caller-saves registers whose values are still needed
2. puts arguments into the build area at the top of the current frame
3. places the static link (if any) in register `ecx`
4. executes a `call` instruction

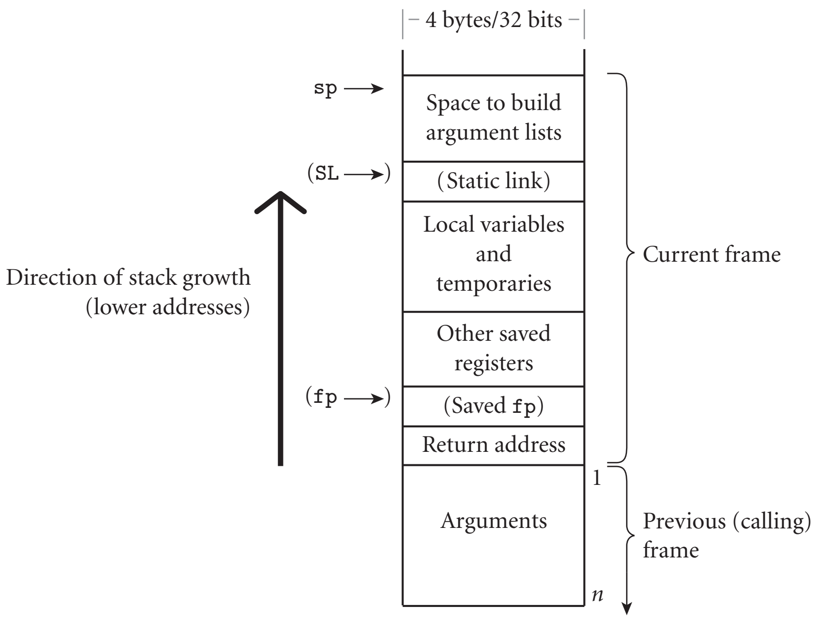


Figure 9.9 **Layout of the subroutine call stack for the GNU Compiler Collection (`gcc`) on 32-bit x86.** The return address is present in all frames. All other parts of the frame are optional; they are present only if required by the current subroutine. In x86 terminology, the `sp` is named `esp`; the `fp` is `ebp` (extended base pointer). The static link, in languages with nested subroutines, is passed in register `ecx`. `SL` marks the location that will be referenced by the static link (if any) of any subroutine nested immediately inside this one. A routine that is neither innermost nor outermost will save its own static link at the location referenced by the static link of its children.

The caller-saves registers consist of `eax`, `edx`, and `ecx`. Step 1 is skipped if none of these contain a value that will be needed later. Step 2 is skipped if the subroutine has no parameters. Step 3 is skipped if the language has no nested subroutines, or if the called routine is declared at the outermost nesting level. The `call` instruction pushes the return address and jumps to the subroutine.

In its prologue, the callee

1. pushes the `fp` onto the stack (if the current routine uses the `fp`), implicitly decrementing the `sp` by 4 (one word).
2. copies the `sp` into the `fp` if necessary, thereby establishing a frame pointer for the current routine
3. pushes any callee-saves registers whose values may be overwritten by the current routine
4. pushes the static link (`ecx`) if the language has nested subroutines and this is not a leaf
5. subtracts the remainder of the frame size from the `sp`

The callee-saves registers are `ebx`, `esi`, `edi`, and, for routines that don't need a frame pointer, `ebp`. For routines that do need a frame pointer, registers `esp` and `ebp` (the `sp` and `fp`, respectively) are saved by Steps 1 and 2. The instructions for some of these steps may be replaced with equivalent sequences by the compiler's code improver, and may be mixed into the rest of the subroutine by the instruction scheduler. In particular, if the value subtracted from the `sp` in Step 5 is made large enough to accommodate the callee-saves registers, then the `pushes` in Steps 3 and 4 may be moved after Step 5 and replaced with `fp`- or `sp`-relative stores.

In its epilogue, the callee

1. sets the return value
2. restores any callee-saved registers
3. copies the `fp` into the `sp`, or subtracts a constant from the `sp`, as appropriate, thereby deallocating the frame
4. pops the `fp`, if any, off the stack
5. returns

Steps 3 and 4 may be effected on the x86 by a single `leave` instruction. As in the previous case study, the caller moves the return value, if it is in a register, to wherever it is needed. It restores any caller-saves registers lazily over time. ■

EXAMPLE 9.61
Subroutine closure trampoline

Because Ada allows subroutines to nest (and Ada 2005 allows arbitrary subroutines to be passed as parameters), a subroutine *S* that is passed as a parameter from *P* to *Q* must be represented by a closure, as described in Section 3.6.1. In many compilers the closure is a data structure containing the address of *S* and the static link that should be used when *S* is called. In `gcc`, however, the closure contains an *x86 code sequence* known as a *trampoline*—typically a pair of instructions to load `ecx` with the appropriate static link and then jump to the beginning of *S*. The trampoline resides in the "local variables and temporaries" section of *P*'s activation record. Its address is passed to *Q*. Rather than "interpret" the closure at run time, *Q* actually `calls` it. One advantage of this mechanism is its interoperability across programming languages: C functions passed as parameters are simply code addresses. In fact, if *S* is declared at the outermost level of lexical nesting, then `gcc` can pass an ordinary code address even when compiling Ada source; in this case no trampoline is required. ■

x86-64 As noted in Section C-5.4.5, the x86-64 has 16 integer registers instead of only 8. AMD, which developed the ISA for the wider architecture, suggests a calling sequence that makes more use of registers (and less of the stack), in a manner reminiscent of ARM (Example C-9.58) and other RISC machines. The GNU compiler generally conforms to AMD's suggestions.

Figure C-9.10 shows a stack frame for the x86-64. The first six integer arguments are passed in registers `rdi`, `rsi`, `rdx`, `rcx`, `r8`, and `r9`, in that order. The static link, when needed, is passed in `r10` (not `rcx`). Registers `rbx` and `r12`–`r15`

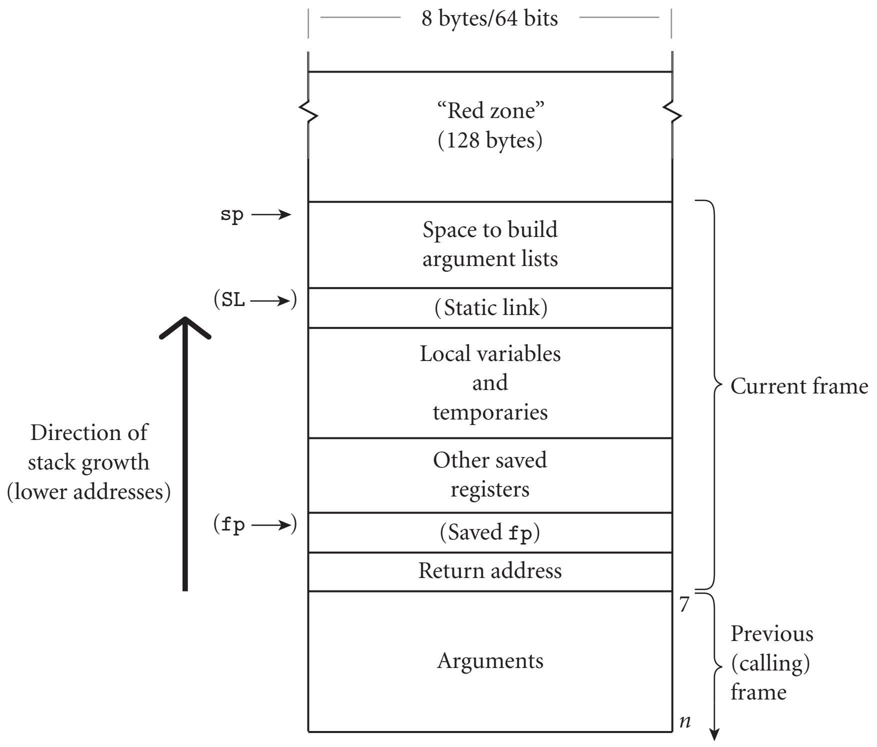


Figure 9.10 Layout of the subroutine call stack for the GNU Compiler Collection (gcc) on 64-bit x86. Conventions differ from those of Figure C-9.9 in three principal ways: (1) most data are 64 bits wide; (2) the first 6 integer arguments are passed in registers rather than on the stack; (3) leaf routines are permitted to use up to 128 bytes of space beyond the top of the stack, without updating the sp.

are callee saves; rax, r10, r11, and the argument registers are caller-saves. Integer function values are returned in rax and (if needed) rdx. The first eight floating-point arguments are passed in XMM/SSE registers xmm0–xmm7 (the legacy x87 registers are for the most part ignored). Additional floating-point arguments are passed on the stack. Floating-point function values are returned in xmm0 and (if needed) xmm1. The stack is always 16-byte aligned at the time of a call.

EXAMPLE 9.62
The x86-64 red zone

Perhaps the most interesting difference between the x86-32 and x86-64 conventions is AMD's specification of a "red zone" beyond the sp. Where the last used word on the stack is guaranteed on x86-32 to be at an address no lower than the sp, on x86-64 it can be up to 128 bytes beyond this point—in effect, the sp protects not only the data at higher addresses (below it in the stack), but up to 128 bytes of additional data as well. Signal handlers and other system software are required to respect this convention. As a result, leaf routines that need a stack frame smaller than 128 bytes need not update the sp. For frequent calls to very small routines, the two-instruction savings in per-call bookkeeping can be significant. ■

✓ **CHECK YOUR UNDERSTANDING**

49. For each of our three case studies, explain which aspects of the calling sequence and stack layout are dictated by the hardware, and which are a matter of software convention.
50. Why don't LLVM and `gcc` restore caller-saves registers immediately after a call?
51. What is a subroutine closure *trampoline*? How does it differ from the usual implementation of a closure described in Section 3.6.1? What are the comparative advantages of the two alternatives?
52. Explain the circumstances under which a subroutine needs a frame pointer (i.e., under which access via displacement addressing from the stack pointer will not suffice).
53. Under what circumstances must an argument that was passed in a register also be saved into the stack?
54. What is the purpose of the "red zone" on x86-64?

DESIGN & IMPLEMENTATION

9.10 Executing code in the stack

A disadvantage of trampoline-based closures is the need to execute code in the stack. Many machines and operating systems disallow such execution, for at least two important reasons. First, as noted in Section C-5.1, modern microprocessors typically have separate instruction and data caches, for fast concurrent access. Allowing a process to write and execute the same region of memory means that these caches must be kept mutually consistent (coherent), a task that introduces significant hardware complexity (on some machines it requires execution of a special hardware instruction). Second, many computer security breaches involve a *code injection* attack, in which an intruder exploits software vulnerabilities (e.g., the lack of array bounds checking in C) to write instructions into the stack, and to overwrite the saved return address so that execution will jump into that code when the current subroutine returns. Such an attack is possible only on machines in which writable data are also executable. When compiling code for use on modern systems, `gcc` embeds a call to a library routine that reverses the system default and re-enables stack execution prior to using a trampoline.

9 Subroutines and Control Abstraction

9.2.3 Register Windows

EXAMPLE 9.63

Register windows on the SPARC

As an alternative to saving and restoring registers on subroutine calls and returns, the original Berkeley RISC machines [PD80, Pat85] incorporated a hardware mechanism known as *register windows*. The basic idea is to provide a very large set of physical registers, most of which are organized as a collection of overlapping windows (Figure C-9.11). A few register names (`r0`–`r7` in the figure) always refer to the same locations, but the rest (`r8`–`r31` in the figure) are interpreted relative to the currently active window. On a subroutine call, the hardware moves to a different window. To facilitate the passing of parameters, the old and new windows overlap: the top few registers in the caller's window (`r24`–`r31` in the figure) are the same as the bottom few registers in the callee's window (`r8`–`r15` in the figure). On a machine with register windows, the compiler places values of use only within the current subroutine in the middle part of the window. It copies values to the upper part of the window to pass them to a called routine, within which they are read from the lower part of the window. ■

Since the number of physical windows is fixed, a long chain of subroutine calls can cause the hardware to run off the end of the register set, resulting in a "window overflow" interrupt that drops the processor into the operating system. The interrupt handler then treats the set of available windows as a circular buffer. It copies the contents of one or more windows to memory and then resumes execution. Later, a "window underflow" interrupt will occur when control attempts to return into a window whose contents have been written to memory. Again the operating system recovers, by restoring the saved registers and resuming execution. In practice, eight windows appear to suffice to make overflow and underflow relatively rare in typical programs.

Register windows have been used in several RISC processors, but only one of these, the SPARC, is commercially significant today. The Intel IA-64 (Itanium), introduced shortly after the turn of the century, also uses register windows, though it is not a RISC machine. The advantage of windows, of course, is

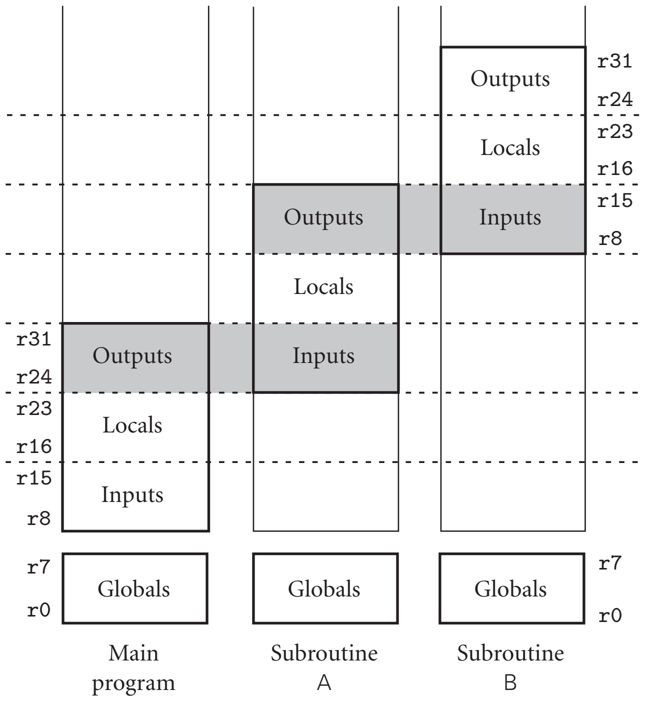


Figure 9.11 **Register windows.** When the main program calls subroutine A, and again when A calls B, register names `r0`–`r7` continue to refer to the same locations, but register names `r8`–`r31` are changed to refer to a new, overlapping window. High-numbered registers in the caller share locations with low-numbered registers in the callee.

that they reduce the number of loads and stores required for the typical subroutine call. At the same time, register windows significantly increase the amount of state associated with the currently running program. When the operating system decides to give the processor to a different application for a while (something that most systems do many times per second), it must save all this state to memory, or arrange for the processor to trap back into the OS if the new process attempts to access an unsaved window. Worse, while register windows nicely capture the referencing environment of a single thread of control, they do not work well for languages that need more than one referencing environment (execution context). Several language features, including continuations (Section 6.2.2), iterators (Section 6.5.3), and coroutines (Section 9.5), are difficult to implement on a machine with register windows, because they require that we save and restore not only the visible registers, but those in other windows as well, when switching between contexts. It is unclear whether the reduction in subroutine call overhead outweighs the extra cost of context switches for typical application workloads, particularly given that loads and stores for parameters are almost always cache hits.

✓ CHECK YOUR UNDERSTANDING

55. What are *register windows*? What purpose do they serve?
56. Which commercial instruction sets include register windows?
57. Explain the concepts of register window *overflow* and *underflow*.
58. Why are register windows a potential problem for multithreaded programs?

9 Subroutines and Control Abstraction

9.3.2 Call by Name

Call by name implements the normal-order argument evaluation described in Section 6.6.2. A call-by-name parameter is reevaluated in the caller's referencing environment every time it is used. The effect is as if the called routine had been textually expanded at the point of call, with the actual parameter (which may be a complicated expression) replacing every occurrence of the formal parameter. To avoid the usual problems with macro parameters, the "expansion" is defined to include parentheses around the replaced parameter wherever syntactically valid, and to make "suitable systematic changes" to the names of any formal parameters or local identifiers that share the same name, so that their meanings never conflict [NBB$^+$63, p. 12]. Call by name was the default in Algol 60; call by value was available as an alternative. In Simula call by value was the default; call by name was the alternative.

To implement call by name, Algol 60 implementations passed a hidden subroutine that evaluated the actual parameter in the caller's referencing environment. Such a hidden routine is usually called a *thunk*.[1] In most cases thunks are trivial. If an actual parameter is a variable name, for example, the thunk simply reads the variable from memory. In some cases, however, a thunk can be elaborate. Perhaps the most famous occurs in what is known as *Jensen's device*, named after Jørn Jensen [Rut67]. The idea is to pass to a subroutine both a built-up expression and one or more of the variables used in the expression. Then by changing the values of the individual variable(s), the called routine can deliberately and systematically change the value of the built-up expression. This device can be used, for example, to write a summation routine:

EXAMPLE 9.64
Jensen's device

1 In general, a thunk is a procedure of zero arguments used to delay evaluation of an expression. Other examples of thunks can be seen in the `delay` mechanism of Example 6.88 and the `promise` constructor of Exercise 11.18.

```
real procedure sum(expr, i, low, high);
    value low, high;
        comment low and high are passed by value;
        comment expr and i are passed by name;
    real expr;
    integer i, low, high;
begin
    real rtn;
    rtn := 0;
    for i := low step 1 until high do
        rtn := rtn + expr;
        comment the value of expr depends on the value of i;
    sum := rtn
end sum
```

Now to evaluate the sum

$$y = \sum_{1 \le x \le 10} 3x^2 - 5x + 2$$

we can simply say

```
y := sum(3*x*x - 5*x + 2, x, 1, 10);
```

Label Parameters

Both Algol 60 and Algol 68 allowed a label to be passed as a parameter. If a called routine performed a `goto` to such a label, control would usually need to escape the local context, unwinding the subroutine call stack as it did so. Details of the unwinding operation would depend on the location of the label. For each intervening scope, the `goto` would have to restore saved registers, deallocate the

DESIGN & IMPLEMENTATION

9.11 Call by name

In practice, most uses of call by name in Algol 60 and Simula programs served to allow a subroutine to change the value of an actual parameter; neither language offered call by reference. Unfortunately, call by name is significantly more expensive than call by reference: it requires the invocation of a thunk (as opposed to a simple indirection) on every use of a formal parameter. Call by name is also prone to subtle program bugs when a change to a variable in a surrounding scope unintentionally alters the value of a formal parameter. (Call by reference suffers from a milder form of this problem, as discussed in Example 3.20.) Such deliberate subtleties as Jensen's device are comparatively rare, and can be imitated in other languages through the use of formal subroutines. Call by name was dropped in Algol 68, in favor of call by reference.

stack frame, and perform any other operations normally handled by epilogue code. To implement label parameters, Algol implementations typically passed a thunk that performed the appropriate operations for the given label. Note that the target label would generally need to lie in some surrounding scope, where it was visible to the caller under static scoping rules.

Label parameters were usually used to handle *exceptional conditions*—conditions that prevent a subroutine from performing its usual operation, and that cannot be handled in the local context. Instead of returning, an Algol routine that encountered a problem (e.g., invalid input) could perform a `goto` to a label parameter, on the assumption that the label refered to code that would perform some remedial operation, or print an appropriate error message. In more recent languages, label parameters have been replaced by more structured exception handling mechanisms, as discussed in Section 9.4.

CHECK YOUR UNDERSTANDING

59. What is *call by name*? What language first provided it? Why isn't it used by the language's descendants?
60. What is *call by need*? How does it differ from call by name? What modern languages use it?
61. How does a subroutine with call-by-name parameters differ from a macro?
62. What is a *thunk*? What is it used for?
63. What is *Jensen's device*?

DESIGN & IMPLEMENTATION

9.12 Call by need

Functional languages like Miranda and Haskell typically pass parameters using a *memoizing* implementation of normal-order evaluation, as described in Section 6.6.2. This *lazy* implementation is sometimes called *call by need*. Memoization calculates and records the value of a parameter the first time it is needed, and uses the recorded value thereafter. In the absence of side effects, call by need is indistinguishable from call by name. It avoids the expense of repeated evaluation, but precludes the use of techniques like Jensen's device in languages that *do* have side effects. Among imperative languages, call by need appears in the scripting language R, where it serves to avoid the expense of evaluating (even once) any complex arguments that are not actually needed.

9 Subroutines and Control Abstraction

9.5.3 Implementation of Iterators

EXAMPLE 9.65
Coroutine-based iterator invocation

Consider the following `for` loop from Example 6.66:

```
for i in range(first, last, step):
    ...
```

Using coroutines, a compiler might translate this as

```
iter := new from_to_by(first, last, step, i, done, current_coroutine)
while not done do
    . . .
    transfer(iter)
destroy(iter)
```

After the loop completes, the implementation can reclaim the space consumed by iter. ■

EXAMPLE 9.66
Coroutine-based iterator implementation

The definition of from_to_by itself is quite straightforward:

```
coroutine from_to_by(from_val, to_val, by_amt : int;
                         ref i : int; ref done : bool; caller : coroutine)
    i := from_val
    if by_amt > 0 then
        done := from_val ≥ to_val
        detach
        loop
            i +:= by_amt
            done := i ≥ to_val
            transfer(caller)      –– yield i
    else
        done := from_val ≤ to_val
        detach
        loop
            i +:= by_amt
            done := i ≤ to_val
            transfer(caller)      −− yield i
```

Parameters i and done are passed by reference so that the iterator can modify them in the caller's context. The caller's identity is passed as a final argument so that the iterator can tell which coroutine to resume when it has computed the next loop index. Because the caller is named explicitly, it is easy for iterators to nest, as in Figure 6.5. ■

Single-Stack Implementation

While coroutines *suffice* for the implementation of iterators, they are not *necessary*. A simpler, single-stack implementation is also possible. Because a given iterator (e.g., an instance of `from_to_by`) is always resumed at the same place in the code (between iterations of a given `for` loop), we can be sure that the subroutine call stack will always contain the same frames whenever the iterator runs. Moreover, since `yield` statements can appear only in the main body of the iterator (never in nested routines), we can be sure that the stack will always contain the same frames whenever the iterator transfers back to its caller. These two facts imply that we can place the frame of the iterator directly on top of the frame of its caller in a single central stack.

When an iterator is created, its frame is pushed on the stack. When it yields a value, control returns to the `for` loop, but the iterator's frame is left on the stack. If the body of the loop makes any subroutine calls, the frames for those calls will be allocated beyond the frame of the iterator. Since control must return to the loop before the iterator resumes, we know that such frames will be gone again before the iterator has a chance to see them: if it needs to call subroutines itself, the stack above it will be clear. Likewise, if the iterator calls any subroutines, they will return (popping their frames from the stack) before the `for` loop runs again. Nested iterators present no special problems (see Exercise C-9.34).

Data Structure Implementation

EXAMPLE 9.67
Iterator usage in C#

Compilers for C# 2.0 employ yet another implementation of iterators. Like Java, C# 1.1 provided iterator objects. Each such object implements the `IEnumerator` interface, which provides `MoveNext` and `Current` methods. Typically an iterator is obtained by calling the `GetEnumerator` method of an object (a container) that implements the `IEnumerable` interface:

```
for (IEnumerator i = myTree.GetEnumerator(); i.MoveNext();) {
    object o = i.Current;
    Console.WriteLine(o.ToString());
}
```
■

EXAMPLE 9.68
Implementation of C# iterators

C# 2.0 provides true iterators as an extension of iterator objects. The programmer simply declares a method that contains one or more `yield return` statements, and whose return type is `IEnumerator` or `IEnumerable`. Here is an example of the latter:

```
static IEnumerable FromToBy (int fromVal, int toVal, int byAmt)
{
    if (byAmt >= 0) {
        for (int i = fromVal; i <= toVal; i += byAmt) {
            yield return i;
        }
    } else {
        for (int i = fromVal; i >= toVal; i += byAmt) {
            yield return i;
        }
    }
}
```

The compiler automatically transforms this code into a hidden class with a `GetEnumerator` method, along the lines of Figure C-9.12. Within this code, an explicit state variable keeps track of the "program counter" of the last `yield` statement. In addition, local variable `i` of the true iterator becomes a data member of the `FromToByImpl` class, leaving the iterator with no need for a stack frame across iterations of the loop. In a quite literal sense, the compiler transforms each true iterator into an iterator object. ■

Recursive iterators present no particular difficulties: a nested iterator is allocated on demand when the outer iterator enters a `foreach` loop, and is referred to by a reference in that outer iterator. The details are deferred to Exercise C-9.35. Because iterator objects are allocated from the heap, the C# implementation of true iterators may be somewhat slower than the stack-based implementation of the previous subsection.

✓ CHECK YOUR UNDERSTANDING

64. Describe the "obvious" implementation of iterators using coroutines.

65. Explain how the state of multiple active iterators can be maintained in a single stack.

66. Describe the transformation used by C# compilers to turn a true iterator into an iterator object.

```
static IEnumerable FromToBy(int fromVal, int toVal, int byAmt) {
    return new FromToByImpl(fromVal, toVal, byAmt);
}
class FromToByImpl : IEnumerator, IEnumerable {
    enum State {starting, goingUp, goingDown, done}
    int i, tv, ba;
    State s;

    public FromToByImpl(int fromVal, int toVal, int byAmt) {
        i = fromVal; tv = toVal; ba = byAmt; s = State.starting;
    }
    public IEnumerator GetEnumerator() {
        return this;
    }
    public object Current {
        get { return i; }
    }
    public bool MoveNext() {
        switch (s) {
            case State.starting :
                if (ba >= 0) {
                    if (i <= tv) { s = State.goingUp; return true; }
                    else { s = State.done; return false; }
                } else {
                    if (i >= tv) { s = State.goingDown; return true; }
                    else { s = State.done; return false; }
                }
            case State.goingUp :
                i += ba;
                if (i <= tv) return true;
                else { s = State.done; return false; }
            case State.goingDown :
                i += ba;
                if (i >= tv) return true;
                else { s = State.done; return false; }
            default: // for completeness
            case State.done : return false;
        }
    }
    public void Reset() {
        s = State.starting;
    }
}
```

Figure 9.12 Iterator object equivalent of a true iterator in C#. This handwritten code corresponds to Example C-9.68. It represents, at the source level, what the compiler creates at the level of intermediate code: a state machine that tracks the program counter of the original iterator, with a starting state, an ending state, and one state for each `yield return` statement. The arms of the `switch` statement capture the code paths in the original iterator that move from one state to the next.

9 Subroutines and Control Abstraction

9.5.4 Discrete Event Simulation

EXAMPLE 9.69

Sequential simulation of a complex physical system

Suppose that we wish to experiment with the flow of traffic in a city. A computerized traffic model, if it captures the real world with sufficient accuracy, will allow us to predict the effects of construction projects, accidents, increased traffic due to new development, or changes to the layout of streets. It is difficult (though certainly not impossible) to write such a simulation in a conventional sequential language. We would probably represent each interesting object (automobile, intersection, street segment, etc.) with a data structure. Our main program would then look something like this:

```
while current_time < end_of_simulation
    calculate next time t at which an interesting interaction will occur
    current_time := t
    update state of objects to reflect the interaction
    record desired statistics
print collected statistics
```

The problem with this approach lies in determining which objects will interact next, and in remembering their state from one interaction to the next. It is in some sense unnatural to represent active objects such as cars with passive data structures, and to make time the active entity in the program. An arguably more attractive approach is to represent each active object with a coroutine, and to let each object keep track of its own state.

EXAMPLE 9.70

Initialization of a coroutine-based traffic simulation

If each active object can tell when it will next do something interesting, then we can determine which objects will interact next by keeping the currently inactive coroutines in a priority queue, ordered by the time of their next event. We might begin a one-day traffic simulation by creating a coroutine for each trip to be taken by a car that day, and inserting each coroutine into the priority queue with a "wakeup" time indicating when the trip is to begin:

```
coroutine trip(. . . )
. . .
for each trip t
    p := new trip(. . . )
    schedule(p, t.start_time)
```

EXAMPLE 9.71
Traversing a street segment in the traffic simulation

Let us assume that we think of street segments as passive, and represent them with data structures. At any given moment, we can model a segment by the number of cars that it is carrying in each direction. This number in turn will affect the speed at which the cars can safely travel. Whenever it awakens, the coroutine representing a trip examines the next street segment over which it needs to travel. Based on the current load on that segment, it calculates how much time it will take to traverse it, and schedules itself to awaken again at an appropriate point in the future:

```
coroutine trip(origin, destination : location)
    plan a route from origin to destination
    detach
    for each segment of the route
        calculate time $i$ to reach the end of the segment
        schedule(current_coroutine, current_time + $i$)
```

EXAMPLE 9.72
Scheduling a coroutine for future execution

The `schedule` operation is easily built on top of transfer:

```
schedule(p : coroutine; t : time)
    –– p may be self or other
    insert (p, t) in priority queue
    if p = current_coroutine      –– self
        extract earliest pair ($q$, $s$) from priority queue
        current_time := $s$
        transfer($q$)
```

EXAMPLE 9.73
Queueing cars at a traffic light

In some cases, it may be difficult to determine when to reschedule a given object. Suppose, for example, that we wish to more accurately model the effects of traffic signals at intersections. We might represent each traffic signal with a data structure that records the waiting cars in each direction, and a coroutine that lets cars through as the signal changes color:

```
record controlled_intersection =
    EW_cars, NS_cars : queue of trip
    const per_car_lag_time : time
        –– how long it takes a car to start after its predecessor does
    coroutine signal(EW_duration, NS_duration : time)
        detach
        loop
            change_time := current_time + EW_duration
            while current_time < change_time
                if EW_cars not empty
                    schedule(dequeue(EW_cars), current_time)
                schedule(current_coroutine, current_time + per_car_lag_time)
```

```
change_time := current_time + NS_duration
while current_time < change_time
    if NS_cars not empty
        schedule(NS_cars.dequeue(), current_time)
    schedule(current_coroutine, current_time + per_car_lag_time)
```

EXAMPLE 9.74
Waiting at a light

When it reaches the end of a street segment that is controlled by a traffic signal, a trip need not calculate how long it will take to get through the intersection. Rather, it enters itself into the appropriate queue of waiting cars and "goes to sleep," knowing that the `signal` coroutine will awaken it at some point in the future:

```
coroutine trip(origin, destination : location)
    plan a route from origin to destination
    detach
    for each segment of the route
        calculate time i to reach the end of the segment
        schedule(current_coroutine, current_time + i)
        if end of segment has a traffic light
            identify appropriate queue Q
            Q.enqueue(current_coroutine)
            sleep()
```

EXAMPLE 9.75
Sleeping in anticipation of future execution

Like `schedule`, `sleep` is easily built on top of transfer:

```
sleep()
    extract earliest pair (q, s) from priority queue
    current_time := s
    transfer(q)
```

The `schedule` operation, in fact, is simply

```
schedule(p : coroutine; t : time)
    insert (p, t) in priority queue
    if p = current_coroutine
        sleep()
```

Obviously this traffic simulation is too simplistic to capture the behavior of cars in a real city, but it illustrates the basic concepts of discrete event simulation. More sophisticated simulations are used in a wide range of application domains, including all branches of engineering, computational biology, physics and cosmology, and even computer design. Multiprocessor simulations (see reference [VF94], for example) are typically divided into a "front end" that simulates the processors and a "back end" that simulates the memory subsystem. Each coroutine in the front end consists of a machine-language interpreter that captures the behavior of one of the system's processing cores. Each coroutine in the back end represents a load or a store instruction. Every time a processor performs a load or store, the front end creates a new coroutine in the back end. Data structures in the back end represent various hardware resources, including caches, buses, network links, message routers, and memory modules. The coroutine for a given load or store checks to see if its location is in the local cache. If not, it must traverse the interconnection network between the processor and memory, competing with other coroutines for access to hardware resources, much as cars in our simple example compete for access to street segments and intersections. The behavior of the back-end system in turn affects the front end, since a processor must wait for a load to complete before it can use the data, and since the rate at which stores can be injected into the back end is limited by the rate at which they propagate to memory.

CHECK YOUR UNDERSTANDING

67. Summarize the computational model of discrete event simulation. Explain the significance of the time-based priority queue.
68. When building a discrete event simulation, how does one decide which things to model with coroutines, and which to model with data structures?
69. Are all inactive coroutines guaranteed to be in the priority queue? Explain.

9 Subroutines and Control Abstraction

9.8 Exercises

9.26 Suppose you wish to minimize the size of closures in a language implementation that uses a display to access nonlocal objects. Assuming a language like Pascal or Ada, in which subroutines have limited extent, explain how an appropriate display for a formal subroutine can be calculated when that routine is finally called, starting with only (1) the value of the frame pointer, saved in the closure at the time that the closure was created, (2) the subroutine return addresses found in the stack at the time the formal subroutine is finally called, and (3) static tables created by the compiler. How costly is your scheme?

9.27 Elaborate on the reasons why even parameters passed in registers may sometimes need to have locations in the stack. Consider all the cases in which it may not suffice to keep a parameter in a register throughout its lifetime.

9.28 Most versions of the C library include a function, `alloca`, that dynamically allocates space within the current stack frame.[2] It has two advantages over the usual `malloc`, which allocates space in the heap: it's usually very fast, and the space it allocates is reclaimed automatically when the current subroutine returns. Assuming the programmer *wants* deallocation to happen then, it's convenient to be able to skip the explicit `free` operations. How might you implement `alloca` in conjunction with the calling conventions of our various case studies?

9.29 Explain how to extend the conventions of Figure C-9.9 and Section C-9.2.2 to accommodate arrays whose bounds are not known until elaboration time (as discussed in Section 8.2.2). What ramifications does this have for the use of separate stack and frame pointers?

2 Unfortunately, `alloca` is not POSIX compliant, and implementations vary greatly in their semantics and even in details of the interface. Portable programs are wise to avoid this routine.

9.30 In all three of our case studies, stack-based arguments were placed into the argument build area in "reverse" order, with the lowest-numbered argument at the top. Explain why this is important. (Hint: Consider subroutines with variable numbers of arguments, as discussed in Section 9.3.3.)

9.31 How would you implement nested subroutines as parameters on a machine that doesn't let you execute code in the stack? Can you pass a simple code address, or do you require that closures be interpreted at run time?

9.32 If you have read the rest of Chapter 9, you may have noticed that the term "trampoline" is also used in conjunction with the implementation of signal handlers (Section 9.6.1). What is the connection (if any) between these uses of the term?

9.33 Explain how you might implement `setjmp` and `longjmp` on a SPARC.

9.34 Following the code in Figure 6.5, and assuming a single-stack implementation of iterators, trace the contents of the stack during the execution of a `for` loop that iterates over all nodes of a complete, 3-level (6-node) binary tree.

9.35 Build a preorder iterator for binary trees in Java, C#, or Python. Do not use a true iterator or an explicit stack of tree nodes. Rather, create nested iterator objects on demand, linking them together as a C# compiler might if it were building the iterator object equivalent of a true preorder iterator.

9.36 One source of inaccuracy in the traffic simulation of Section C-9.5.4 has to do with the timing at traffic signals. If a signal is currently green in the `EW` direction, but the queue of waiting cars is empty, the `signal` coroutine will go to sleep until `current_time + EW_duration`. If a car arrives before the coroutine wakes up again, it will needlessly wait. Discuss how you might remedy this problem.

9 Subroutines and Control Abstraction

9.9 Explorations

9.47 Read the ARM calling sequence standard for 64-bit (v8) code. Compare and contrast to the conventions of Section C-9.2.2. Pay particular attention to the lists of caller- and callee-saves registers, and to the registers used to pass arguments. Speculate as to reasons for the differences.

9.48 Research the full range of hardware support for subroutines on the x86, including all variants of `call`. Note that the `leave` instruction is sometimes generated by modern compilers, but others, including `enter`, `pushad`, `popad`, `pushfd`, and `popfd`, usually are not. In addition, the optional argument of `ret` is almost never used, and `push` and `pop` are used sparingly. Discuss the technological trends that have made this machinery obsolete.

9.49 As an example of hard-core CISC design, research the subroutine calling conventions of the Digital VAX. Be sure to describe the behavior of the `calls` instruction in detail.

9.50 Study the implementation of a user-level thread management package written for the SPARC. How does it manage register windows?

9.51 Learn how parameter passing is implemented in the Glasgow Haskell compiler. How expensive is its call-by-need–based lazy evaluation?

9.52 Learn about the Time Warp system for discrete event simulation, developed by David Jefferson and colleagues [JBW$^+$87]. Discuss its relationship to both the classic discrete event simulation of Section C-9.5.4 and the speculative parallelism of mechanisms like transactional memory (to be discussed in Section 13.4.4).

10

Data Abstraction and Object Orientation

10.6 True Multiple Inheritance

EXAMPLE 10.57
Deriving from two base classes (reprise)

Recall our administrative computing example in C++:

```
class student : public person, public system_user { ...
```

To implement multiple inheritance, we must be able to generate both a "`person` view" and a "`system_user` view" of a `student` object on demand, for example when assigning a reference to a `student` object into a `person` or `system_user` variable. For one of the base classes (`person`, say) we can do the same thing we did with single inheritance: let the data members of that base class lie at the beginning of the representation of the derived class, and let the virtual methods of that base class lie at the beginning of the vtable. Then when we assign a reference to a `student` object into a `person` variable, code that manipulates the `person` variable will just use a prefix of the data members and the vtable. ■

EXAMPLE 10.58
(Nonrepeated) multiple inheritance

For the other base class (`system_user`), things get more complicated: we can't put *both* base classes at the beginning of the derived class. One possible solution is shown in Figure C-10.9. It is based loosely on the implementation described by Ellis and Stroustrup [ES90, Chap. 10], and builds on the implementation of mix-ins described in Section 10.5. Because the `system_user` fields of a `student` follow the `person` fields, the assignment of a reference to a `student` object into a variable of type `system_user*` requires that we adjust our "view" by adding the compile-time constant offset *d*.

The vtable for a `student` is broken into two parts. The first part lists the virtual methods of the derived class and the first base class (`person`). The second part lists the virtual methods of the second base class. (We have already introduced a method, `print_mailing_label`, defined in class `person`. We may similarly imagine that `system_user` defines a virtual method `print_stats` that is supposed to dump account statistics to standard output.) Generalization to three or more base classes is straightforward; see Exercise C-10.23.

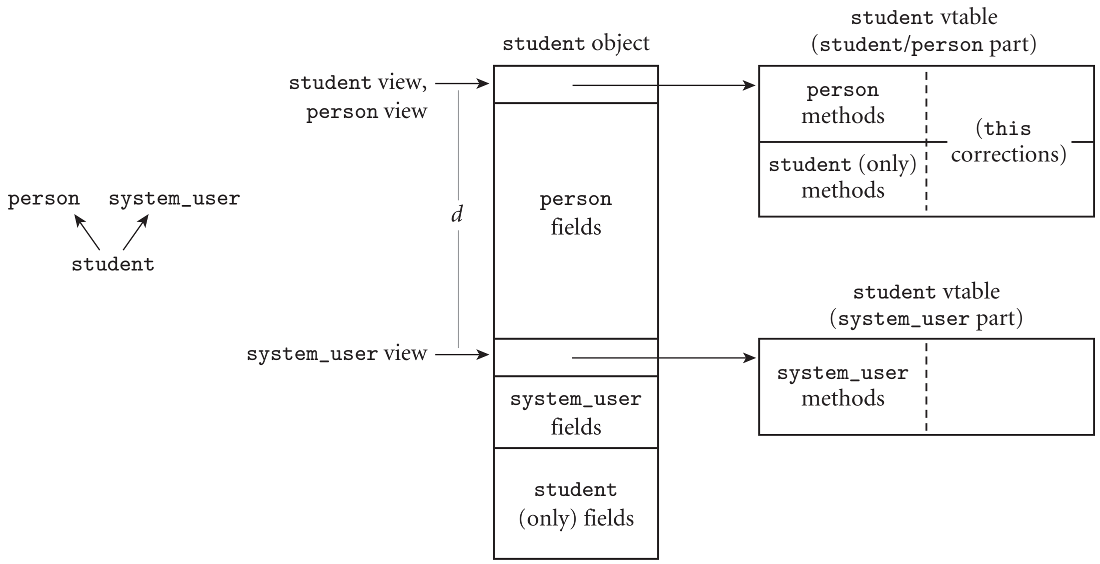


Figure 10.9 Implementation of (nonrepeated) multiple inheritance. The size *d* of the person portion of the object is a compile-time constant. We access the system_user portion of the vtable by adding *d* to the address of a student object before indirecting. Likewise, we create a system_user view of a student object by adding *d* to the object's address. Each vtable entry consists of both a method address and a "this correction" value equal to the signed distance between the view through which the vtable was accessed and the view of the class in which the method was defined.

Every data member of a student object has a compile-time-constant offset from the beginning of the object. Likewise, every virtual method has a compile-time-constant offset from the beginning of one of the parts of the vtable. The address of the person/student portion of the vtable is stored in the beginning of the object. The address of the system_user portion of the vtable is stored at offset *d*. Note that both parts of the vtable are specific to class student. In particular, the system_user part of the vtable is *not* shared by objects of class system_user, because the contents of the tables will be different if student has overridden any of system_user's virtual methods. ■

EXAMPLE 10.59
Method invocation with multiple inheritance

To call the virtual method print_mailing_label, originally defined in person, we can use a code sequence similar to the one shown in Section 10.4.3 for single inheritance. To call a virtual method originally defined in system_user, we must first add the offset *d* to our object's address, in order to find the address of the system_user portion of the vtable. Then we can index into this system_user vtable to find the address of the appropriate method to call. But we are left with one final problem: what is the appropriate value of this to pass to the method?

As a concrete example, suppose that student does not override print_stats (though it certainly could). If our object is of class student, we should pass a system_user view of it to print_stats: the address of the object, plus *d*.

If, however, our object is of some class (transfer_student, perhaps) that does override print_stats, then we should pass a transfer_student view to print_stats. If we are accessing our object through a variable (a reference or a pointer) whose methods are dynamically bound, then we can't tell at compile time which one of these cases applies. Worse yet, we may not even know how to generate a transfer_student view if we have to: class transfer_student may not have been invented when this part of our code was compiled, so we certainly don't know how far into it the system_user fields appear! ■

EXAMPLE 10.60
this correction

A common solution is for each vtable entry to consist of a *pair* of fields. One is the address of the method's code; the other is a "this correction" value, to be added to the view through which we found the vtable. (Compare this to the mix-in approach of Figure 10.7, where a single this correction would suffice for the entire vtable, because we knew that every method would be provided by the class that implements the interface, not by the interface itself.) Returning to Figure C-10.9, the "this correction" field of the vtable entry for print_stats would contain $-d$ if print_stats was overridden by student, and zero otherwise. In the system_user part of the vtable for the (yet to be written) class transfer_student, the "this correction" field might contain some other value $-e$. In general, the "this correction" is the distance between the view of the class in which the method was *declared* (and through which we accessed the vtable) and the view of the class in which the method was *defined* (and which will therefore be expected by the subroutine's implementation).

If variable my_student contains a reference to (a student view of) some object at run time, and if print_stats is the third virtual method of system_user, then the code to call my_student.debug_print would look something like this:

```
r1 := my_student                 -- student view of object
r1 := r1 + d                     -- system_user view of object
r2 := *r1                        -- address of appropriate vtable
r3 := *(r2 + (3−1) × 8)          -- method address
r2 := *(r2 + (3−1) × 8 + 4)      -- this correction
r1 := r1 + r2                    -- this
call *r3
```

Here we have assumed that both method addresses and this corrections are four bytes long, that this is to be passed in r1, and that there are no other arguments. On a typical machine this code is three instructions (including one memory access) longer than the code required with single inheritance, and five instructions (including three memory accesses) longer than a call to a statically identified method. ■

10.6.1 Semantic Ambiguities

In addition to implementation complexities (only some of which we have discussed so far), multiple inheritance introduces potential semantic problems.

EXAMPLE 10.61
Methods found in more than one base class

Suppose that both system_user and person define a print_stats method. If we have a variable s of type student* and we call s->print_stats, which version of the method should we get? In CLOS and Python, we get the version from the base class that appeared first in the derived class's header. In Eiffel, we get a static semantic error if we try to define a derived class with such an ambiguity. In C++, we can define the derived class, but we get a static semantic error if we attempt to use a member whose name is ambiguous. To resolve the ambiguity, we can use the feature renaming mechanism in Eiffel to give different names to the inherited methods. In C++ we must redefine the conflicting method explicitly:

```
void student::print_stats() {
    person::print_stats();
    system_user::print_stats();
}
```

Here we have chosen to call the print_stats routines of both base classes, using the :: scope resolution operator to name them. We could of course have chosen to call just one, or to write our own code from scratch. We could even arrange for access to both routines by giving them new names:

```
void student::print_person_stats() {
    person::print_stats();
}
void student::print_user_stats() {
    system_user::print_stats();
}
```

■

EXAMPLE 10.62
Overriding an ambiguous method

Things are a little messier if either or both of the identically named base class methods are virtual, and we want to override them in the derived class. Following Stroustrup [Str13, Sec. 21.3.3], we can solve the problem by interposing an intermediate class between each base class and the derived class:

DESIGN & IMPLEMENTATION

10.8 The cost of multiple inheritance

The implementation we have described for multiple inheritance, using this corrections in vtables, has the unfortunate property of increasing the overhead of all virtual method invocations, even in programs that do not make use of multiple inheritance. This sort of mandatory overhead is something that language designers (and the designers of systems languages in particular) generally try to avoid; as a matter of principle, complex special cases should not reduce the efficiency of the simpler common case. Fortunately, there are other implementations of multiple inheritance (see Exercise C-10.28) in which the cost of modifying this is paid only when the correction is nonzero.

```
class person_interface : public person {
public:
    virtual void print_person_stats() = 0;
    void print_stats() { print_person_stats(); }
        // overrides person::print_stats
};
class system_user_interface : public system_user {
public:
    virtual void print_user_stats() = 0;
    void print_stats() { print_user_stats(); }
        // overrides system_user::print_stats
};
class student : public person_interface, public system_user_interface {
public:
    void print_person_stats() { ...
    void print_user_stats() { ...
    ...
};
```

We leave it as an exercise (C-10.24) to show what happens if we assign a `student` object into a variable `p` of type `person*` and then call `p->print_stats()`. ■

A more serious ambiguity arises when a class *D* inherits from two base classes, *B* and *C*, both of which inherit from some common base class *A*. In this situation, should an object of class *D* contain one instance of the data members of class *A* or two? The answer would seem to be program dependent. For example, suppose that professors, like students, are all given accounts in our administrative computing system. Then, like class `student`, we might want class `professor` to inherit from both `person` and `system_user`:

EXAMPLE 10.63
Repeated multiple inheritance

```
class professor : public person, public system_user { ...
```

But now suppose that some professors take courses on occasion as nonmatriculated students. In this case we might want a new class that supports both sets of operations:

```
class student_prof : public student, public professor { ...
```

Class `student_prof` inherits from `person` and from `system_user` twice, once each through `student` and `professor`. If we think about it, we probably want a `student_prof` to have *one* instance of the data members of class `person`—one name, one university ID number, one mailing address—and *two* instances of the data members of class `system_user`—separate user accounts (with separate user ids, disk quotas, etc.) for the student and professor roles:

$$\begin{tikzcd}
\texttt{system\_user} & & \texttt{person} & & \texttt{system\_user} \\
& \texttt{student} \arrow[ul] \arrow[ur] & & \texttt{professor} \arrow[ul] \arrow[ur] & \\
& & \texttt{student\_prof} \arrow[ul] \arrow[ur] & &
\end{tikzcd}$$

The `system_user` case—separate copies from each branch of the inheritance tree—is known as *replicated inheritance*. The `person` case—a single copy from both branches of the tree—is known as *shared* inheritance. Both are forms of *repeated inheritance*. ■

EXAMPLE 10.64
Shared inheritance in C++

Replicated inheritance is the default in C++. Shared inheritance is the default in Eiffel. Shared inheritance can be obtained in C++ by specifying that a base class is `virtual`:

```
class student : public virtual person, public system_user { ...
class professor : public virtual person, public system_user { ...
```

In this case the members of class `person` are shared when inherited over multiple paths, while the members of class `system_user` are replicated. ■

EXAMPLE 10.65
Replicated inheritance in Eiffel

Replicated inheritance of individual features can be obtained in Eiffel through the renaming mechanism described in Section 10.2.2:

```
class student inherit person; system_user ...
class professor inherit person; system_user ...

class student_prof
inherit
    student
        rename
            user_id as student_user_id,
                disk_quota as student_disk_quota
            end;
        professor
            rename
                user_id as prof_user_id,
                disk_quota as prof_disk_quota
        end
feature
    ...
end -- class student_prof
```

Features inherited with different final names are replicated; features inherited with the same final name are shared. Multiple inheritance in CLOS is always shared, unless the user interposes interface classes as shown above explicitly; there is no other renaming mechanism. ■

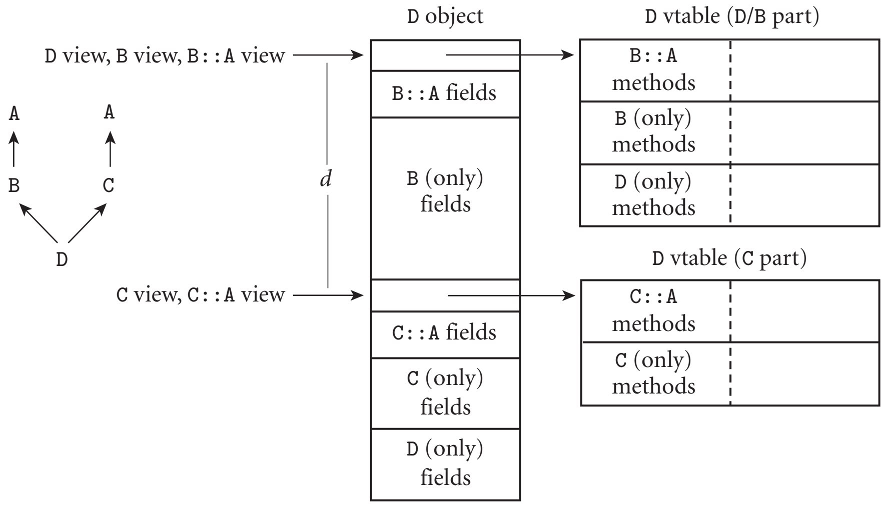


Figure 10.10 **Implementation of replicated multiple inheritance.** Each base class contains a complete copy of class A. As in Figure C-10.9, the vtable for class D is split into two parts, one for each base class, and each vtable entry consists of a ⟨method address, this correction⟩ pair.

10.6.2 Replicated Inheritance

EXAMPLE 10.66
Using replicated inheritance

Replicated inheritance introduces no serious implementation problems beyond those of nonrepeated multiple inheritance. As shown in Figure C-10.10, an object (in this case of class D) that inherits a base class (A) over two different paths in the inheritance tree has two copies of A's data members in its representation, and a set of entries for the virtual methods of A in each of the parts of its vtable. Creation of a B view of a D object (e.g., when assigning a pointer to a D object into a B* variable) would not require the execution of any code. Creation of a C view (e.g., when assigning into a C* variable) would require the addition of offset *d*.

Because of ambiguity, we cannot access A members of a D object by name. We can access them, however, if we assign a pointer to a D object into a B* or C* variable. Similarly, a pointer to a D object cannot be assigned into an A pointer directly: there would be no basis on which to choose the A for which to create a view. We can, however, perform the assignment through a B* or C* intermediary:

```
class A { ...
class B : public A { ...
class C : public A { ...
class D : public B, public C { ...
...
A* a;   B* b;   C* c;   D* d;
a = d;  // error; ambiguous
```

```
b = d;  // ok
c = d;  // ok
a = b;  // ok; a := d's B's A
a = c;  // ok; a := d's C's A
```

As described in Example C-10.60, vtable entries will need to consist of ⟨method address, `this` correction⟩ pairs. ■

10.6.3 Shared Inheritance

EXAMPLE 10.67 Overriding methods with shared inheritance

Shared inheritance introduces a new opportunity for ambiguity and additional implementation complexity. As in the previous subsection, assume that `D` inherits from `B` and `C`, both of which inherit from `A`. This time, however, assume that `A` is shared:

```
class A {
public:
    virtual void f();
    ...
};
class B : public virtual A { ...
class C : public virtual A { ...
class D : public B, public C { ...
```

The new ambiguity arises if `B` or `C` overrides method `f`, declared in `A`: which version (if any) does `D` inherit? C++ defines a reference to `f` to be unambiguous (and therefore valid) if one of the possible definitions *dominates* the others, in the sense that its class is a descendant of the classes of all the other definitions. In our specific example, `D` can inherit an overridden version of `f` from either `B` or `C`. If both of them override it, however, any attempt to use `f` from within `D`'s code will be a static semantic error. Eiffel provides comparatively elaborate mechanisms for controlling ambiguity. A class that inherits an overridden method over more than one path can specify the version it wants. Alternatively, through renaming, it can retain access to all versions. ■

EXAMPLE 10.68 Implementation of shared inheritance

To implement shared inheritance we must recognize that because a single instance of `A` is a part of both `B` and `C`, we cannot make the representations of both `B` and `C` contiguous in memory. In Figure C-10.11, in fact, we have chosen to make neither `B` nor `C` contiguous. We insist, however, that the representation of every `B`, `C`, or `D` object (and every `B`, `C`, or `D` view of an object of a derived class) contain the address of the `A` part of the object at a compile-time constant offset from the beginning of the view. To access a data member of `A`, we first indirect through this address, and then apply the offset of the member within `A`. To call the *n*th virtual method declared in `A`, we execute the following code:

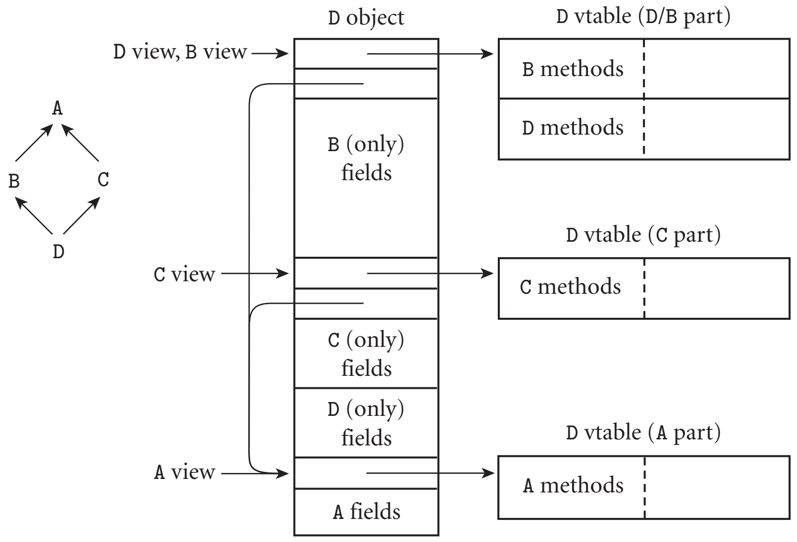


Figure 10.11 Implementation of shared multiple inheritance. Objects of class `B`, `C`, and `D` contain the address of their `A` components at a compile-time constant offset (in this case, immediately after the vtable address). As in Figures C-10.9 and C-10.10, `this` corrections for virtual methods in vtable entries are relative to the view of the class in which the method was declared (i.e., through which the vtable was accessed).

```
r1 := my_D_view                 –– original view of object
r1 := ∗(r1 + 4)                 –– A view
r2 := ∗r1                       –– address of A part of vtable
r3 := ∗(r2 + (n−1) × 8)         –– method address
r2 := ∗(r2 + (n−1) × 8 + 4)     –– this correction
r1 := r1 + r2                   –– this
call ∗r3
```

This code sequence is the same number of instructions in length as our sequence for nonvirtual base classes (Example C-10.60), but involves one more memory access (to indirect through the `A` address). The code will work with any `D` view of any object, including an object of a class derived from `D`, in which the `D` and `A` views might be more widely separated. The constant 4 in the second line assumes 4-byte addresses, with the address of `D`'s `A` part located immediately after `D`'s initial vtable address. In an object with more than one virtual base class, the address of the part of the object corresponding to each such base would be found at a different offset from the beginning of the object. ■

The implementation strategy of Figure C-10.11 works in C++ because we always know when a base class is `virtual` (shared). For data members and virtual methods of nonvirtual base classes, we continue to use the (cheaper) lookup algorithms of Figures C-10.9 and C-10.10. In Eiffel, on the other hand, a feature that is inherited via replication at one level of the class hierarchy may be inherited via sharing later on. As a result, Eiffel requires a somewhat more elaborate implementation strategy (see Exercise C-10.29).

We can avoid the extra level of indirection when accessing virtual methods of virtual base classes in C++ if we are willing to replicate portions of a class's vtable. We explore this option in Exercise C-10.30.

✓ **CHECK YOUR UNDERSTANDING**

48. Give a few examples of the semantic ambiguities that arise when a class has more than one base class.

49. Explain the distinction between replicated and shared multiple inheritance. When is each desirable?

50. Explain how even nonrepeated multiple inheritance introduces the need for "`this` correction" fields in individual vtable entries.

51. Explain how shared multiple inheritance introduces the need for an additional level of indirection when accessing fields of certain parent classes.

52. Explain why true multiple inheritance is harder to implement than mix-in inheritance.

10 Data Abstraction and Object Orientation

10.7.1 The Object Model of Smalltalk

Smalltalk is heavily integrated into its programming environment. In fact, unlike all of the other languages mentioned in this book, a Smalltalk program does not consist of a simple sequence of characters. Rather, Smalltalk programs are meant to be viewed within the *browser* of a Smalltalk implementation, where font changes and screen position can be used to differentiate among various parts of a given program unit. Together with the contemporaneous Interlisp and Pilot/Mesa projects at PARC, the Smalltalk group shares credit for developing the now ubiquitous concepts of bit-mapped screens, windows, menus, and mice.

Smalltalk uses an untyped reference model for all variables. Every variable refers to an object, but the class of the object need not be statically known. As described in Section 10.3.1, every Smalltalk object is an instance of a class descended from a single base class named `Object`. All data are contained in objects. The most trivial of these are simple immutable objects such as `true` (of class `Boolean`) and `3` (of class `Integer`).

EXAMPLE 10.69
Operations as messages in Smalltalk

Operations are all conceptualized as *messages* sent to objects. The expression `3 + 4`, for example, indicates sending a `+` message to the (immutable) object `3`, with a reference to the object `4` as argument. In response to this message, the object `3` creates and returns a reference to the (immutable) object `7`. Similarly, the expression `a + b`, where `a` and `b` are variables, indicates sending a `+` message to the object referred to by `a`, with the reference in `b` as argument. If `a` happens to refer to `3` and `b` refers to `4`, the effect will be the same as it was in the case of the constants. ■

EXAMPLE 10.70
Mixfix messages

As described in Section 6.1, multiargument messages have multiword ("mixfix") names. Each word ends with a colon; each argument follows a word. The expression

```
    myBox displayOn: myScreen at: location
```

sends a `displayOn: at:` message to the object referred to by variable `myBox`, with the objects referred to by `myScreen` and `location` as arguments. ■

EXAMPLE 10.71
Selection as an `ifTrue: ifFalse:` message

Even control flow in Smalltalk is conceptualized as messages. Consider the selection construct:

```
n < 0
    ifTrue: [abs <- n negated]
    ifFalse: [abs <- n]
```

This code begins by sending a `< 0` message (a `<` message with `0` as argument) to the object referred to by `n`. In response to this message, the object referred to by `n` will return a reference to one of two immutable objects: `true` or `false`. This reference becomes the value of the `n < 0` expression.

Smalltalk evaluates expressions left-to-right without precedence or associativity. The value of `n < 0` therefore becomes the recipient of an `ifTrue: ifFalse:` message. This message has two arguments, each of which is a *block*. A block in Smalltalk is a fragment of code enclosed in brackets. It is an immutable object, with semantics roughly comparable to those of a lambda expression in Lisp. To execute a block we send it a `value` message.

When sent an `ifTrue: ifFalse:` message, the immutable object `true` sends a `value` message to its first argument (which had better be a block) and then returns the result. The object `false`, on the other hand, in response to the same message, sends a `value` message to its second argument (the block that followed `ifFalse:`). The left arrow (`<-`) in each block is the assignment operator. Assignment is not a message; it is a side effect of evaluation of the right-hand side. As in expression-based languages such as Algol 68, the value of an assignment expression is the value of the right-hand side. The overall value of our selection expression will be the value of one of the blocks, namely a reference to `n` or to its additive inverse, whichever is non-negative. For the sake of convenience, Boolean objects in Smalltalk also implement `ifTrue:`, `ifFalse:`, and `ifFalse: ifTrue:` methods. ■

EXAMPLE 10.72
Iterating with messages

Iteration is modeled in a similar fashion. For enumeration-controlled loops, class `Integer` implements `timesRepeat:` and `to: by: do:` methods:

```
pow <- 1.
10 timesRepeat:
    [pow <- pow * n]

sum <- 0.
1 to: 100 by: 2 do:
    [:i | sum <- sum + (a at: i)]
```

The first of these code fragments calculates $\mathtt{n}^{10}$. In response to a `timesRepeat:` message, the integer *k* sends a `value` message to the argument (a block) *k* times. The second code fragment sums the odd-indexed elements of the array referred to by `a`. In response to a `to: by: do:` message, the integer *k* behaves as one might expect: it sends a `value:` message to its third argument (a block) $\lfloor(t-k+b)/b\rfloor$ times, where t is the first argument and b is the second argument. Note the colon at the end of `value:`. The plain `value` message is unary; the `value:` message has an argument; it is understood by blocks that have a (single) formal parameter. In our loop example, the integer `1` sends the messages `value: 1`, `value: 3`, `value: 5`, and so on to the block `[:i | sum <- sum + (a at: i)]`. The `:i |` at the beginning of the block is its formal parameter. The `at:` message is understood by arrays. For iteration with a step size of one, integers also provide a `to: do:` method. ■

EXAMPLE 10.73
Blocks as closures

Because it is an object, a block can be referred to by a variable:

```
b <- [n <- n + 1].            " b is now a closure"
c <- [:i | n <- n + i].       " so is c"
...
b value.                      " increment n by 1"
c value: 3.                   " increment n by 3"
```

A block with two parameters expects a `value: value:` message. A block with j parameters expects a message whose name consists of the word `value:` repeated j times. Comments in Smalltalk are double-quoted (strings are single-quoted). ■

EXAMPLE 10.74
Logical looping with messages

For logically controlled loops, Smalltalk relies on the `whileTrue:` message, understood by blocks:

```
tail <- myList.
[tail next ~~ nil]
    whileTrue: [tail <- tail next]
```

This code sets `tail` to the final element of `myList`. The double-tilde (`~~`) operator means "does not refer to the same object as." The method `next` is assumed to return a reference to the element following its recipient. In response to a `whileTrue:` message, a block sends itself a `value` message. If the result of that message is a reference to `true`, the block sends a `value` message to the argument of the original message and repeats. Blocks also implement a `whileFalse:` method. ■

EXAMPLE 10.75
Defining control abstractions

The blocks of Smalltalk allow the programmer to construct almost arbitrary control-flow constructs. Because of their simple syntax, Smalltalk blocks are even easier to manipulate than the lambda expressions of Lisp. In effect, a `to: by: do:` message turns iteration "inside out," making the body of the loop a simple message argument that can be executed (by sending it a `value` message) from within the body of the `to: by: do:` method. Smalltalk programmers can define similar methods for other container classes, obtaining all the power of iterators (Section 6.5.3) and much of the power of `call_with_current_continuation` (Section 9.4.3):

```
myTree inorderDo: [:node | whatever ]
```
■

It is worth noting that the uniform object model of computation in Smalltalk does not necessarily imply a uniform implementation. Just as Clu implementations implement built-in immutable objects as values, despite their reference semantics (Section 6.1.2), a Smalltalk implementation is likely to use the usual machine instructions for computer arithmetic, rather than actually sending messages to integers. In a similar vein, the most common control-flow constructs (`ifTrue: ifFalse:`, `to: by: do:`, `whileTrue:`, etc.) are likely to be recognized by a Smalltalk interpreter, and implemented with special, faster code.

EXAMPLE 10.76
Recursion in Smalltalk

We end this subsection by observing that recursion works at least as well in Smalltalk as it does in other imperative languages. The following is a recursive implementation of Euclid's algorithm:

```
gcd: other                                          "other is a formal parameter"
    (self = other)
        ifTrue:   [↑self].                          "end condition"
    (self < other)
        ifTrue:   [↑self gcd: (other - self)]               "recurse"
        ifFalse:  [↑other gcd: (self - other)]              "recurse"
```

The up-arrow (↑) symbol is comparable to the `return` of C or Algol 68. The keyword `self` is comparable to `this` in C++. We have shown the code in mixed fonts, much as it would appear in a Smalltalk browser. The header of the method is identified by bold face type. ■

✓ CHECK YOUR UNDERSTANDING

53. Name the three projects at Xerox PARC in the 1970s that pioneered modern GUI-based personal computers.
54. Explain the concept of a *message* in Smalltalk.
55. How does Smalltalk indicate multiple message arguments?
56. What is a *block* in Smalltalk? What mechanism does it resemble in Lisp?
57. Give three examples of how Smalltalk models control flow as message evaluation.
58. Explain how type checking works in Smalltalk.

10 Data Abstraction and Object Orientation

10.9 Exercises

10.23 Suppose that class D inherits from classes A, B, and C, none of which share any common ancestor. Show how the data members and vtable(s) of D might be laid out in memory. Also show how to convert a reference to a D object into a reference to an A, B, or C object.

10.24 Consider the person_interface and system_user_interface classes described in Example C-10.62. If student is derived from person_interface and system_user_interface, explain what happens in the following method call:

```
student s;
person *p = &s;
...
p->print_stats();
```

You may wish to use a diagram of the representation of a student object to illustrate the method lookups that occur and the views that are computed. You may assume an implementation akin to that of Figure C-10.10, without shared inheritance.

10.25 Given the inheritance tree of Example C-10.63, show a representation for objects of class student_prof. You may want to consult Figures C-10.9, C-10.10, and C-10.11.

10.26 Given the memory layout of Figure C-10.9 and the following declarations:

```
student& sr;
system_user& ur;
```

show the code that must be generated for the assignment

```
ur = sr;
```

(Pitfall: Be sure to consider null pointers.)

10.27 Standard C++ provides a "pointer-to-member" mechanism for classes:

```
class C {
public:
    int a;
    int b;
} c;
int C::*pm = &C::a;
    // pm points to member a of an (arbitrary) C object
...
C* p = &c;
p->*pm = 3;     // assign 3 into c.a
```

Pointers to members are also permitted for subroutine members (methods), including virtual methods. How would you implement pointers to virtual methods in the presence of C++-style multiple inheritance?

10.28 As an alternative to using ⟨method address, `this` correction⟩ pairs in the vtable entries of a language with multiple inheritance, we could leave the entries as simple pointers, but make them point to code that updates `this` in-line, and then jumps to the beginning of the appropriate method, much as Java 8 and Scala do to implement default methods on a standard Java virtual machine. Show the sequence of instructions executed under this scheme. What factors will influence whether it runs faster or slower than the sequence shown in Example C-10.60? Which scheme will use less space? (Remember to count both code and data structure size, and consider which instructions must be replicated at every call site.)

Pursuing the replacement of data structures with executable code even further, consider an implementation in which the vtable itself consists of executable code. Show what this code would look like and, again, discuss the implications for time and space overhead.

10.29 In Eiffel, shared inheritance is the default rather than the exception. Only renamed features are replicated. As a result, it is not possible to tell when looking at a class whether its members will be inherited replicated or shared by derived classes. Describe a uniform mechanism for looking up members inherited from base classes that will work whether they are replicated *or* shared. (Hint: Consider the use of dope vectors for records containing arrays of dynamic shape, as described in Section 8.2.2. For further details, consult the compiler text of Wilhelm and Maurer [WM95, Sec. 5.3].)

10.30 In Figure C-10.11, consider calls to virtual methods declared in `A`, but called through a `B`, `C`, or `D` object view. We could avoid one level of indirection by appending a copy of the `A` part of the vtable to the `D/B` and

`C` parts of the vtable (with suitably adjusted `this` corrections). Give calling sequences for this alternative implementation. In the worst case, how much larger may the vtable be for a class with *n* ancestors?

10.31 Consider the Smalltalk implementation of Euclid's algorithm, presented at the end of Section C-10.7.1. Trace the messages involved in evaluating `4 gcd: 6`.

10

Data Abstraction and Object Orientation

10.10 Explorations

10.38 Figure out how multiple inheritance is implemented in your local C++ compiler. How closely does it follow the strategy of Sections C-10.6.2 and C-10.6.3? What rationale do you see for any differences?

10.39 Learn how multiple inheritance is implemented in Perl and Python (you might begin by reading Section 14.4.4). Describe the differences with respect to Sections C-10.6.2 and C-10.6.3. Discuss the advantages and drawbacks of dynamic typing in object-oriented languages.

11 Functional Languages

11.7 Theoretical Foundations

EXAMPLE 11.77

Functions as mappings

Mathematically, a function is a single-valued mapping: it associates every element in one set (the *domain*) with (at most) one element in another set (the *range*). In conventional notation, we indicate the domain and range by writing

$$\text{sqrt} : \mathcal{R} \longrightarrow \mathcal{R}$$

We can, of course, have functions of more than one variable—that is, functions whose domains are Cartesian products:

$$\text{plus} : [\mathcal{R} \times \mathcal{R}] \longrightarrow \mathcal{R}$$

■

If a function provides a mapping for every element of the domain, the function is said to be *total*. Otherwise, it is said to be *partial*. Our *sqrt* function is partial: it does not provide a mapping for negative numbers. We could change our definition to make the domain of the function the non-negative numbers, but such changes are often inconvenient, or even impossible: inconvenient because we should like all mathematical functions to operate on $\mathcal{R}$; impossible because we may not know which elements of the domain have mappings and which do not. Consider for example the function f that maps every natural number a to the smallest natural number b such that the digits of the decimal representation of a appear b digits to the right of the decimal point in the decimal expansion of π. Clearly $f(59) = 4$, because $\pi = 3.14159\ldots$. But what about $f(428945028)$, or in general $f(n)$ for arbitrary n? Absent results from number theory, it is not at all clear how to characterize the values at which f is defined. In such a case a partial function is essential.

EXAMPLE 11.78

Functions as sets

It is often useful to characterize functions as sets or, more precisely, as subsets of the Cartesian product of the domain and the range:

$$\text{sqrt} \subset [\mathcal{R} \times \mathcal{R}]$$
$$\text{plus} \subset [\mathcal{R} \times \mathcal{R} \times \mathcal{R}]$$

We can specify *which* subset using traditional set notation:

$$\text{sqrt} \equiv \left\{(x, y) \in \mathcal{R} \times \mathcal{R} \mid y > 0 \wedge x = y^2\right\}$$
$$\text{plus} \equiv \{(x, y, z) \in \mathcal{R} \times \mathcal{R} \times \mathcal{R} \mid z = x + y\}$$

Note that this sort of definition tells us what the value of a function like sqrt is, but it does *not* tell us how to compute it; more on this distinction below. ■

EXAMPLE 11.79

Functions as powerset elements

One of the nice things about the set-based characterization is that it makes it clear that a function is an ordinary mathematical object. We know that a function from A to B is a subset of $A \times B$. This means that it is an *element* of the *powerset* of $A \times B$—the set of all subsets of $A \times B$, denoted $2^{A \times B}$:

$$\text{sqrt} \in 2^{\mathcal{R} \times \mathcal{R}}$$

Similarly,

$$\text{plus} \in 2^{\mathcal{R} \times \mathcal{R} \times \mathcal{R}}$$

Note the overloading of notation here. The powerset 2^A should not be confused with exponentiation, though it is true that for a finite set A the number of elements in the powerset of A is 2^n, where $n = |A|$, the cardinality of A. ■

EXAMPLE 11.80

Function spaces

Because functions are single-valued, we know that they constitute only *some* of the elements of $2^{A \times B}$. Specifically, they constitute all and only those sets of pairs in which the first component of each pair is unique. We call the set of such sets the *function space* of A into B, denoted $A \to B$. Note that $(A \to B) \subset 2^{A \times B}$. In our examples:

$$\text{sqrt} \in [\mathcal{R} \to \mathcal{R}]$$
$$\text{plus} \in [(\mathcal{R} \times \mathcal{R}) \to \mathcal{R}]$$

EXAMPLE 11.81

Higher-order functions as sets

Now that functions are elements of sets, we can easily build higher-order functions:

$$\text{compose} \equiv \{(f, g, h) \mid \forall x \in \mathcal{R},\ h(x) = f(g(x))\}$$

What are the domain and range of compose? We know that f, g, and h are elements of $\mathcal{R} \to \mathcal{R}$. Thus

$$\text{compose} \in [(\mathcal{R} \to \mathcal{R}) \times (\mathcal{R} \to \mathcal{R})] \to (\mathcal{R} \to \mathcal{R})$$

Note the similarity to the notation employed by the ML type system (Section 7.2.4). ■

EXAMPLE 11.82

Curried functions as sets

Using the notion of "currying" from Section 11.6, we note that there is an alternative characterization for functions like plus. Rather than a function from pairs of reals to reals, we can capture it as a function from reals to functions from reals to reals:

$$\text{curried\_plus} \in \mathcal{R} \to (\mathcal{R} \to \mathcal{R})$$ ■

We shall have more to say about currying in Section C-11.7.3.

11.7.1 Lambda Calculus

As we suggested in the main text, one of the limitations of the function-as-set notation is that it is *nonconstructive*: it doesn't tell us how to *compute* the value of a function at a given point (i.e., on a given input). Church designed the lambda calculus to address this limitation. In its pure form, lambda calculus represents *everything* as a function. The natural numbers, for example, can be represented by a distinguished zero function (commonly the identity function) and a successor function. (One common formulation uses a select_second function that takes two arguments and returns the second of them. The successor function is then defined in such a way that the number n ends up being represented by a function that, when applied to select_second n times, returns the identity function [Mic89, Sec. 3.5]; [Sta95, Sec. 7.6]; see Exercise C-11.23.) While of theoretical importance, this formulation of arithmetic is highly cumbersome. We will therefore take ordinary arithmetic as a given in the remainder of this subsection. (And of course all practical functional programming languages provide built-in support for both integer and floating-point arithmetic.)

A lambda expression can be defined recursively as (1) a *name*; (2) a lambda *abstraction* consisting of the letter λ, a name, a dot, and a lambda expression; (3) a function *application* consisting of two adjacent lambda expressions; or (4) a parenthesized lambda expression. To accommodate arithmetic, we will extend this definition to allow numeric literals.

EXAMPLE 11.83
Juxtaposition as function application

When two expressions appear adjacent to one another, the first is interpreted as a function to be applied to the second:

sqrt n

Most authors assume that application associates left-to-right (so $f\,A\,B$ is interpreted as $(f\,A)\,B$, rather than $f\,(A\,B)$), and that application has higher precedence than abstraction (so $\lambda x.A\,B$ is interpreted as $\lambda x.(A\,B)$, rather than $(\lambda x.A)\,B$). ML adopts these rules. ■

EXAMPLE 11.84
Lambda calculus syntax

Parentheses are used as necessary to override default groupings. Specifically, if we distinguish between lambda expressions that are used as functions and those that are used as arguments, then the following unambiguous CFG can be used to generate lambda expressions with a minimal number of parentheses:

$$\begin{aligned}
expr &\longrightarrow \texttt{name} \mid number \mid \lambda\ \texttt{name}\ .\ expr \mid\ func\ arg \\
func &\longrightarrow \texttt{name} \mid (\ \lambda\ \texttt{name}\ .\ expr\ ) \mid\ func\ arg \\
arg &\longrightarrow \texttt{name} \mid number \mid (\ \lambda\ \texttt{name}\ .\ expr\ ) \mid (\ func\ arg\ )
\end{aligned}$$

In words: we use parentheses to surround an abstraction that is used as either a function or an argument, and around an application that is used as an argument. ■

EXAMPLE 11.85

Binding parameters with λ

The letter λ introduces the lambda calculus equivalent of a formal parameter. The following lambda expression denotes a function that returns the square of its argument:

$$\lambda x.\mathsf{times}\ x\, x$$

The name (variable) introduced by a λ is said to be *bound* within the expression following the dot. In programming language terms, this expression is the variable's scope. A variable that is not bound is said to be *free*. ■

EXAMPLE 11.86

Free variables

As in a lexically scoped programming language, a free variable needs to be defined in some surrounding scope. Consider, for example, the expression $\lambda x.\lambda y.\mathsf{times}\ x\, y$. In the inner expression ($\lambda y.\mathsf{times}\ x\, y$), y is bound but x is free. There are no restrictions on the use of a bound variable: it can play the role of a function, an argument, or both. Higher-order functions are therefore completely natural. ■

EXAMPLE 11.87

Naming functions for future reference

If we wish to refer to them later, we can give expressions names:

$$\begin{aligned}
\mathsf{square} &\equiv \lambda x.\mathsf{times}\ x\, x\\
\mathsf{identity} &\equiv \lambda x.x\\
\mathsf{const7} &\equiv \lambda x.7\\
\mathsf{hypot} &\equiv \lambda x.\lambda y.\mathsf{sqrt}\ (\mathsf{plus}\ (\mathsf{square}\ x)\ (\mathsf{square}\ y))
\end{aligned}$$

Here $\equiv$ is a metasymbol meaning, roughly, "is an abbreviation for." ■

EXAMPLE 11.88

Evaluation rules

To compute with the lambda calculus, we need rules to evaluate expressions. It turns out that three rules suffice:

beta reduction: For any lambda abstraction $\lambda x.E$ and any expression M, we say

$$(\lambda x.E)\, M \to_\beta E[M\backslash x]$$

where $E[M\backslash x]$ denotes the expression E with all free occurrences of x replaced by M. Beta reduction is not permitted if any free variables in M would become bound in $E[M\backslash x]$.

alpha conversion: For any lambda abstraction $\lambda x.E$ and any variable y that has no free occurrences in E, we say

$$\lambda x.E \to_\alpha \lambda y.E[y\backslash x]$$

eta reduction: A rule to eliminate "surplus" lambda abstractions. For any lambda abstraction $\lambda x.E$, where E is of the form $F\, x$, and x has no free occurrences in F, we say

$$\lambda x.F\, x \to_\eta F$$

■

EXAMPLE 11.89
Delta reduction for arithmetic

To accommodate arithmetic we will also allow an expression of the form op x y, where x and y are numeric literals and op is one of a small set of standard functions, to be replaced by its arithmetic value. This replacement is called *delta reduction*. In our examples we will need only the functions plus, minus, and times:

$$\begin{aligned}\mathsf{plus}\ 2\ 3 &\rightarrow_\delta 5\\ \mathsf{minus}\ 5\ 2 &\rightarrow_\delta 3\\ \mathsf{times}\ 2\ 3 &\rightarrow_\delta 6\end{aligned}$$

■

Beta reduction resembles the use of call by name parameters (Section 9.3.1). Unlike Algol 60, however, the lambda calculus provides no way for an argument to carry its referencing environment with it; hence the requirement that an argument not move a variable into a scope in which its name has a different meaning. Alpha conversion serves to change names to make beta reduction possible. Eta reduction is comparatively less important. If square is defined as above, eta reduction allows us to say that

EXAMPLE 11.90
Eta reduction

$$\lambda x.\mathsf{square}\ x \rightarrow_\eta \mathsf{square}$$

In English, square is a function that squares its argument; $\lambda x.$square x is a function of x that squares x. The latter reminds us explicitly that it's a function (i.e., that it takes an argument), but the former is a little less messy looking. ■

Through repeated application of beta reduction and alpha conversion (and possibly eta reduction), we can attempt to reduce a lambda expression to its simplest possible form—a form in which no further beta reductions are possible. An example can be found in Figure C-11.5. In line (2) of this derivation we have to employ an alpha conversion because the argument that we need to substitute for g contains a free variable (h) that is bound within g's scope. If we were to make the substitution of line (3) without first having renamed the bound h (as k), then the free h would have been *captured*, erroneously changing the meaning of the expression.

EXAMPLE 11.91
Reduction to simplest form

In line (5) of the derivation, we had a choice as to which subexpression to reduce. At that point the expression as a whole consisted of a function application in which the argument was itself a function application. We chose to substitute the main argument $((\lambda x.x\,x)\,(\lambda x.x\,x))$, unevaluated, into the body of the main lambda abstraction. This choice is known as *normal-order* reduction, and corresponds to normal-order evaluation of arguments in programming languages, as discussed in Sections 6.6.2 and 11.5. In general, whenever more than one beta reduction could be made, normal order chooses the one whose λ is left-most in the overall expression. This strategy substitutes arguments into functions before reducing them. The principal alternative, *applicative-order* reduction, reduces both the function part and the argument part of every function application to the simplest possible form before substituting the latter into the former. ■

$$\begin{array}{lll}
 & (\underline{\lambda f}.\lambda g.\lambda h.fg(h\,h))\underline{(\lambda x.\lambda y.x)}h(\lambda x.x\,x) & \\
\rightarrow_\beta & (\lambda g.\underline{\lambda h}.(\lambda x.\lambda y.x)g(\underline{h\,h}))h(\lambda x.x\,x) & (1) \\
\rightarrow_\alpha & (\underline{\lambda g}.\lambda k.(\lambda x.\lambda y.x)g(k\,k))\underline{h}(\lambda x.x\,x) & (2) \\
\rightarrow_\beta & (\underline{\lambda k}.(\lambda x.\lambda y.x)h(k\,k))\underline{(\lambda x.x\,x)} & (3) \\
\rightarrow_\beta & (\underline{\lambda x}.\lambda y.x)\underline{h}((\lambda x.x\,x)\ (\lambda x.x\,x)) & (4) \\
\rightarrow_\beta & (\underline{\lambda y}.h)\underline{((\lambda x.x\,x)\ (\lambda x.x\,x))} & (5) \\
\rightarrow_\beta & h & (6)
\end{array}$$

Figure 11.5 Reduction of a lambda expression. The top line consists of a function applied to three arguments. The first argument (underlined) is the "select first" function, which takes two arguments and returns the first. The second argument is the symbol h, which must be either a constant or a variable bound in some enclosing scope (not shown). The third argument is an "apply to self" function that takes one argument and applies it to itself. The particular series of reductions shown occurs in normal order. It terminates with a simplest (normal) form of simply h.

Church and Rosser showed in 1936 that simplest forms are unique: any series of reductions that terminates in a nonreducible expression will produce the same result. Not all reductions terminate, however. In particular, there are expressions for which no series of reductions will terminate, and there are others in which normal-order reduction will terminate but applicative-order reduction will not. The example expression of Figure C-11.5 leads to an infinite "computation" under applicative-order reduction. To see this, consider the expression at line (5). This line consists of the constant function $(\lambda y.h)$ applied to the argument $(\lambda x.x\,x)\ (\lambda x.x\,x)$. If we attempt to evaluate the argument before substituting it into the function, we run through the following steps:

EXAMPLE 11.92
Nonterminating applicative-order reduction

$$\begin{array}{ll}
 & (\underline{\lambda x}.x\,x)\ \underline{(\lambda x.x\,x)} \\
\rightarrow_\beta & (\underline{\lambda x}.x\,x)\ \underline{(\lambda x.x\,x)} \\
\rightarrow_\beta & (\underline{\lambda x}.x\,x)\ \underline{(\lambda x.x\,x)} \\
\rightarrow_\beta & (\underline{\lambda x}.x\,x)\ \underline{(\lambda x.x\,x)} \\
\dots &
\end{array}$$

■

In addition to showing the uniqueness of simplest (normal) forms, Church and Rosser showed that if any evaluation order will terminate, normal order will. This pair of results is known as the *Church-Rosser theorem*.

11.7.2 Control Flow

We noted at the beginning of the previous subsection that arithmetic can be modeled in the lambda calculus using a distinguished zero function (commonly the identity) and a successor function. What about control-flow constructs—selection and recursion in particular?

EXAMPLE 11.93
Booleans and conditionals

The select_first function, $\lambda x.\lambda y.x$, is commonly used to represent the Boolean value true. The select_second function, $\lambda x.\lambda y.y$, is commonly used to represent the Boolean value false. Let us denote these by T and F. The nice thing about these definitions is that they allow us to define an if function very easily:

$$\text{if} \equiv \lambda c.\lambda t.\lambda e.c\,t\,e$$

Consider:

$$\begin{aligned} \text{if}\ T\ 3\ 4 &\equiv (\lambda c.\lambda t.\lambda e.c\,t\,e)\ (\lambda x.\lambda y.x)\ 3\ 4 \\ &\to_\beta^* (\lambda x.\lambda y.x)\ 3\ 4 \\ &\to_\beta^* 3 \end{aligned}$$

$$\begin{aligned} \text{if}\ F\ 3\ 4 &\equiv (\lambda c.\lambda t.\lambda e.c\,t\,e)\ (\lambda x.\lambda y.y)\ 3\ 4 \\ &\to_\beta^* (\lambda x.\lambda y.y)\ 3\ 4 \\ &\to_\beta^* 4 \end{aligned}$$

■

Functions like equal and greater_than can be defined to take numeric values as arguments, returning T or F.

EXAMPLE 11.94
Beta abstraction for recursion

Recursion is a little tricky. An equation like

$$\begin{aligned} \text{gcd} \equiv\ & \lambda a.\lambda b.(\text{if (equal } a\, b)\ a \\ & \quad (\text{if (greater\_than } a\, b)\ (\text{gcd (minus } a\, b)\ b)\ (\text{gcd (Minus } b\, a)\ a))) \end{aligned}$$

is not really a definition at all, because gcd appears on both sides. Our previous definitions (T, F, if) were simply shorthand: we could substitute them out to obtain a pure lambda expression. If we try that with gcd, the "definition" just gets bigger, with new occurrences of the gcd name. To obtain a real definition, we first rewrite our equation using *beta abstraction* (the opposite of beta reduction):

$$\begin{aligned} \text{gcd} \equiv\ & (\lambda g.\lambda a.\lambda b.(\text{if (equal } a\, b)\ a \\ & \quad (\text{if (greater\_than } a\, b)\ (g(\text{minus } a\, b)\ b)\ (g(\text{minus } b\, a)\ a))))\ \text{gcd} \end{aligned}$$

Now our equation has the form

$$\text{gcd} \equiv f\ \text{gcd}$$

where f is the perfectly well-defined (nonrecursive) lambda expression

$$\begin{aligned} & \lambda g.\lambda a.\lambda b.(\text{if (equal } a\, b)\ a \\ & \quad (\text{if (greater\_than } a\, b)\ (g\ (\text{minus } a\, b)\ b)\ (g\ (\text{minus } b\, a)\ a))) \end{aligned}$$

Clearly gcd is a fixed point of f. ■

EXAMPLE 11.95

The fixed-point combinator **Y**

As it turns out, for any function f given by a lambda expression, we can find the least fixed point of f, if there is one, by applying the *fixed-point combinator*

$$\lambda h.(\lambda x.h(xx))\,(\lambda x.h(xx))$$

commonly denoted **Y**. **Y** has the property that for any lambda expression f, if the normal-order evaluation of $\mathbf{Y}f$ terminates, then $f(\mathbf{Y}f)$ and $\mathbf{Y}f$ will reduce to the same simplest form (see Exercise C-11.21). In the case of our gcd function, we have

$$\begin{aligned}\text{gcd} \;\equiv\;& (\lambda h.(\lambda x.h(x\,x))\,(\lambda x.h(x\,x)))\\ & (\lambda g.\lambda a.\lambda b.(\text{if }(\text{equal }a\,b)\;a\\ &\quad (\text{if }(\text{greater\_than }a\,b)\;(g(\text{minus }a\,b)\;b)\;(g(\text{minus }b\,a)\;a))))\end{aligned}$$

Figure C-11.6 traces the evaluation of gcd 4 2. Given the existence of the **Y** combinator, most authors permit recursive "definitions" of functions, for convenience. ■

11.7.3 Structures

EXAMPLE 11.96

Lambda calculus list operators

Just as we can use functions to build numbers and truth values, we can also use them to encapsulate values in structures. Using Scheme terminology for the sake of clarity, we can define simple list-processing functions as follows:

$$\begin{aligned}\text{cons} &\equiv \lambda a.\lambda d.\lambda x.x\,a\,d\\ \text{car} &\equiv \lambda l.l\ \text{select\_first}\\ \text{cdr} &\equiv \lambda l.l\ \text{select\_second}\\ \text{nil} &\equiv \lambda x.T\\ \text{null?} &\equiv \lambda l.l(\lambda x.\lambda y.F)\end{aligned}$$

where select_first and select_second are the functions $\lambda x.\lambda y.x$ and $\lambda x.\lambda y.y$, respectively—functions we also use to represent true and false. ■

EXAMPLE 11.97

List operator identities

Using these definitions we can see that

$$\begin{aligned}\text{car}(\text{cons }A\,B) &\equiv (\lambda l.l\ \text{select\_first})\ (\text{cons }A\,B)\\ &\to_\beta (\text{cons }A\,B)\ \text{select\_first}\\ &\equiv ((\lambda a.\lambda d.\lambda x.x\,a\,d)\,A\,B)\ \text{select\_first}\\ &\to_\beta^* (\lambda x.x\,A\,B)\ \text{select\_first}\\ &\to_\beta \text{select\_first}\ A\,B\\ &\equiv (\lambda x.\lambda y.x)\,A\,B\\ &\to_\beta^* A\end{aligned}$$

$$\begin{aligned}
\mathsf{gcd}\ 2\,4 &\equiv \mathbf{Y} f\ 2\,4\\
&\equiv ((\lambda h.(\lambda x.h(x\,x))\ (\lambda x.h(x\,x)))f)\ 2\,4\\
&\to_\beta ((\lambda x.f(x\,x))\ (\lambda x.f(x\,x)))\ 2\,4\\
&\equiv (k\,k)\ 2\,4,\quad \text{where}\ k \equiv \lambda x.f(x\,x)\\
&\to_\beta (f(k\,k))\ 2\,4\\
&\equiv ((\lambda g.\lambda a.\lambda b.(\text{if}\ (=a\,b)\ a\ (\text{if}\ (>a\,b)\ (g(-\,a\,b)\ b)\ (g(-\,b\,a)\ a))))\ (k\,k))\ 2\,4\\
&\to_\beta (\lambda a.\lambda b.(\text{if}\ (=a\,b)\ a\ (\text{if}\ (>a\,b)\ ((k\,k)(-\,a\,b)\ b)\ ((k\,k)(-\,b\,a)\ a))))\ 2\,4\\
&\to_\beta^* \text{if}\ (=2\,4)\ 2\ (\text{if}\ (>2\,4)\ ((k\,k)\ (-\,2\,4)\ 4)\ ((k\,k)\ (-\,4\,2)\ 2))\\
&\equiv (\lambda c.\lambda t.\lambda e.c\,t\,e)\ (=2\,4)\ 2\ (\text{if}\ (>2\,4)\ ((k\,k)\ (-\,2\,4)\ 4)\ ((k\,k)\ (-\,4\,2)\ 2))\\
&\to_\beta^* (=2\,4)\ 2\ (\text{if}\ (>2\,4)\ ((k\,k)\ (-\,2\,4)\ 4)\ ((k\,k)\ (-\,4\,2)\ 2))\\
&\to_\delta F\,2\ (\text{if}\ (>2\,4)\ ((k\,k)\ (-\,2\,4)\ 4)\ ((k\,k)\ (-\,4\,2)\ 2))\\
&\equiv (\lambda x.\lambda y.y)\ 2\ (\text{if}\ (>2\,4)\ ((k\,k)\ (-\,2\,4)\ 4)\ ((k\,k)\ (-\,4\,2)\ 2))\\
&\to_\beta^* \text{if}\ (>2\,4)\ ((k\,k)\ (-\,2\,4)\ 4)\ ((k\,k)\ (-\,4\,2)\ 2)\\
&\to \ldots\\
&\to (k\,k)\ (-\,4\,2)\ 2\\
&\equiv ((\lambda x.f(x\,x))k)\ (-\,4\,2)\ 2\\
&\to_\beta (f(k\,k))\ (-\,4\,2)\ 2\\
&\equiv ((\lambda g.\lambda a.\lambda b.(\text{if}\ (=a\,b)\ a\ (\text{if}\ (>a\,b)\ (g(-\,a\,b)\ b)\ (g(-\,b\,a)\ a))))\ (k\,k))\ (-\,4\,2)\ 2\\
&\to_\beta (\lambda a.\lambda b.(\text{if}\ (=a\,b)\ a\ (\text{if}\ (>a\,b)\ ((k\,k)(-\,a\,b)\ b)\ ((k\,k)(-\,b\,a)\ a))))\ (-\,4\,2)\ 2\\
&\to_\beta^* \text{if}\ (=(-\,4\,2)\ 2)\ (-\,4\,2)\ (\text{if}\ (>(-\,4\,2)\ 2)\ ((k\,k)\ (-\,(-\,4\,2)\ 2)\ 2)\ ((k\,k)\ (-\,2\ (-\,4\,2))\ (-\,4\,2)))\\
&\equiv (\lambda c.\lambda t.\lambda e.c\,t\,e)\\
&\ (=(-\,4\,2)\ 2)\ (-\,4\,2)\ (\text{if}\ (>(-\,4\,2)\ 2)\ ((k\,k)\ (-\,(-\,4\,2)\ 2)\ 2)\ ((k\,k)\ (-\,2\ (-\,4\,2))\ (-\,4\,2)))\\
&\to_\beta^* (=(-\,4\,2)\ 2)\ (-\,4\,2)\ (\text{if}\ (>(-\,4\,2)\ 2)\ ((k\,k)\ (-\,(-\,4\,2)\ 2)\ 2)\ ((k\,k)\ (-\,2\ (-\,4\,2))\ (-\,4\,2)))\\
&\to_\delta (=2\,2)\ (-\,4\,2)\ (\text{if}\ (>(-\,4\,2)\ 2)\ ((k\,k)\ (-\,(-\,4\,2)\ 2)\ 2)\ ((k\,k)\ (-\,2\ (-\,4\,2))\ (-\,4\,2)))\\
&\to_\delta T\ (-\,4\,2)\ (\text{if}\ (>(-\,4\,2)\ 2)\ ((k\,k)\ (-\,(-\,4\,2)\ 2)\ 2)\ ((k\,k)\ (-\,2\ (-\,4\,2))\ (-\,4\,2)))\\
&\equiv (\lambda x.\lambda y.x)\ (-\,4\,2)\ (\text{if}\ (>(-\,4\,2)\ 2)\ ((k\,k)\ (-\,(-\,4\,2)\ 2)\ 2)\ ((k\,k)\ (-\,2\ (-\,4\,2))\ (-\,4\,2)))\\
&\to_\beta^* (-\,4\,2)\\
&\to_\delta 2
\end{aligned}$$

Figure 11.6 **Evaluation of a recursive lambda expression.** As explained in the body of the text, gcd is defined to be the fixed-point combinator **Y** applied to a beta abstraction f of the standard recursive definition for greatest common divisor. Specifically, **Y** is $\lambda h.(\lambda x.h(x\,x))\ (\lambda x.h(x\,x))$ and f is $\lambda g.\lambda a.\lambda b.(\text{if}\ (=a\,b)\ a\ (\text{if}\ (>a\,b)\ (g(-\,a\,b)\ b)\ (g(-\,b\,a)\ a)))$. For brevity we have used =, >, and − in place of equal, greater_than, and minus. We have performed the evaluation in normal order.

$$\begin{aligned}
\text{cdr}(\text{cons}\ A\ B) &\equiv (\lambda l.l\ \text{select\_second})\ (\text{cons}\ A\ B)\\
&\to_\beta (\text{cons}\ A\ B)\ \text{select\_second}\\
&\equiv ((\lambda a.\lambda d.\lambda x.x\ a\ d)\ A\ B)\ \text{select\_second}\\
&\to_\beta^* (\lambda x.x\ A\ B)\ \text{select\_second}\\
&\to_\beta \text{select\_second}\ A\ B\\
&\equiv (\lambda x.\lambda y.y)\ A\ B\\
&\to_\beta^* B
\end{aligned}$$

$$\begin{aligned}
\text{null?}\ \text{nil} &\equiv (\lambda l.l\ (\lambda x.\lambda y.\text{select\_second}))\ \text{nil}\\
&\to_\beta \text{nil}\ (\lambda x.\lambda y.\text{select\_second})\\
&\equiv (\lambda x.\text{select\_first})\ (\lambda x.\lambda y.\text{select\_second})\\
&\to_\beta \text{select\_first}\\
&\equiv T
\end{aligned}$$

$$\begin{aligned}
\text{null?}\ (\text{cons}\ A\ B) &\equiv (\lambda l.l\ (\lambda x.\lambda y.\text{select\_second}))\ (\text{cons}\ A\ B)\\
&\to_\beta (\text{cons}\ A\ B)\ (\lambda x.\lambda y.\text{select\_second})\\
&\equiv ((\lambda a.\lambda d.\lambda x.x\ a\ d)\ A\ B)\ (\lambda x.\lambda y.\text{select\_second})\\
&\to_\beta^* (\lambda x.x\ A\ B)\ (\lambda x.\lambda y.\text{select\_second})\\
&\to_\beta (\lambda x.\lambda y.\text{select\_second})\ A\ B\\
&\to_\beta^* \text{select\_second}\\
&\equiv F
\end{aligned}$$

■

EXAMPLE 11.98
Nesting of lambda expressions

Because every lambda abstraction has a single argument, lambda expressions are naturally curried. We generally obtain the effect of a multiargument function by nesting lambda abstractions:

$$\text{compose} \equiv \lambda f.\lambda g.\lambda x.f\ (g\ x)$$

which groups as

$$\lambda f.(\lambda g.(\lambda x.(f\ (g\ x))))$$

We commonly think of compose as a function that takes two functions as arguments and returns a third function as its result. We could just as easily, however, think of compose as a function of three arguments: the f, g, and x above. The official story, or course, is that compose is a function of one argument that evaluates to a function of one argument that in turn evaluates to a function of one argument. ■

EXAMPLE 11.99
Paired arguments and currying

If desired, we can use our structure-building functions to define a noncurried version of compose whose (single) argument is a pair:

$$\text{paired\_compose} \equiv \lambda p.\lambda x.(\text{car } p)\,((\text{cdr } p)\, x)$$

If we consider the pairing of arguments as a general technique, we can write a curry function that reproduces the single-argument version, just as we did in Scheme in Section 11.6:

$$\text{curry} \equiv \lambda f.\lambda a.\lambda b.f\,(\text{cons } a\, b)$$ ■

✓ CHECK YOUR UNDERSTANDING

29. What is the difference between *partial* and *total* functions? Why is the difference important?
30. What is meant by the *function space* $A \rightarrow B$?
31. Define *beta reduction*, *alpha conversion*, *eta reduction*, and *delta reduction*.
32. How does beta reduction in lambda calculus differ from lazy evaluation of arguments in a nonstrict programming language like Haskell?
33. Explain how lambda expressions can be used to represent Boolean values and control flow.
34. What is *beta abstraction*?
35. What is the **Y** combinator? What useful property does it possess?
36. Explain how lambda expressions can be used to represent structured values such as lists.
37. State the *Church-Rosser theorem*.

11 Functional Languages

11.10 Exercises

11.20 In Figure C-11.6 we evaluated our expression in normal order. Did we really have any choice? What would happen if we tried to use applicative order?

11.21 Prove that for any lambda expression f, if the normal-order evaluation of $\mathbf{Y}f$ terminates, where $\mathbf{Y}$ is the fixed-point combinator $\lambda h.(\lambda x.h(x\,x))(\lambda x.h(x\,x))$, then $f(\mathbf{Y}f)$ and $\mathbf{Y}f$ will reduce to the same simplest form.

11.22 Given the definition of structures (lists) in Section C-11.7.3, what happens if we apply car or cdr to nil? How might you introduce the notion of "type error" into lambda calculus?

11.23 Let

$$\begin{aligned} \mathsf{zero} &\equiv \lambda x.x \\ \mathsf{succ} &\equiv \lambda n.(\lambda s.(s\ \mathsf{select\_second})\ n) \end{aligned}$$

where $\mathsf{select\_second} \equiv \lambda x.\lambda y.y$. Now let

$$\begin{aligned} \mathsf{one} &\equiv \mathsf{succ\ zero} \\ \mathsf{two} &\equiv \mathsf{succ\ one} \end{aligned}$$

Show that

$$\begin{aligned} \mathsf{one\ select\_second} &= \mathsf{zero} \\ \mathsf{two\ select\_second\ select\_second} &= \mathsf{zero} \end{aligned}$$

In general, show that

$$\mathsf{succ}^n\ \mathsf{zero}\ \mathsf{select\_second}^n = \mathsf{zero}$$

Use this result to define a predecessor function pred. You may ignore the issue of the predecessor of zero.

Note that our definitions of T and F allow us to check whether a number is equal to zero:

$$\mathsf{iszero} \equiv \lambda n.(n\ \mathsf{select\_first})$$

Using succ, pred, iszero, and if, show how to define plus and times recursively. These definitions could of course be made nonrecursive by means of beta abstraction and **Y**.

11 Functional Languages

11.11 Explorations

11.30 Learn about the *typed lambda calculus*. What properties does it have that standard lambda calculus does not? What restrictions does it place on permissible expressions? Possible places to start include Cardelli and Wegner's classic survey [CW85] or the newer text by Pierce [Pie02].

11.31 Learn more about *fixed points*. We mentioned these when presenting the **Y** combinator in Section C-11.7.2. They also arise in the denotational definition of loop constructs, in metacircular interpreters [AS96, Sec. 4.1]), and in the *data flow analysis* used by optimizing compilers (Section C-17.4.2). What do these subjects have in common? Are there important differences as well?

11.32 Explore the connection between lexical scoping in Scheme or OCaml and the notion of free and bound variables in lambda calculus. How closely are these related? Why does lambda calculus require alpha conversion but Scheme and OCaml do not? Is there any analogy in lambda calculus to the dynamic scoping of early dialects of Lisp?

12 Logic Languages

12.3 Theoretical Foundations

In mathematical logic, a *predicate* is a function that maps constants (atoms) or variables to the values true and false. *Predicate calculus* provides a notation and inference rules for constructing and reasoning about *propositions* (*statements*) composed of predicate applications, *operators*, and the *quantifiers* $\forall$ and $\exists$.[1] Operators include and ($\wedge$), or ($\vee$), not ($\neg$), implication ($\rightarrow$), and equivalence ($\leftrightarrow$). Quantifiers are used to introduce bound variables in an appended proposition, much as λ introduces variables in the lambda calculus. The *universal* quantifier, $\forall$, indicates that the proposition is true for all values of the variable. The *existential* quantifier, $\exists$, indicates that the proposition is true for at least one value of the variable. Here are a few examples:

EXAMPLE 12.39
Propositions

$$\forall C[\text{rainy}(C) \wedge \text{cold}(C) \rightarrow \text{snowy}(C)]$$

(For all cities C, if C is rainy and C is cold, then C is snowy.)

$$\forall A, \forall B[(\exists C[\text{takes}(A, C) \wedge \text{takes}(B, C)]) \rightarrow \text{classmates}(A, B)]$$

(For all students A and B, if there exists a class C such that A takes C and B takes C, then A and B are classmates.)

$$\forall N[(N > 2) \rightarrow \neg(\exists A, \exists B, \exists C[A^N + B^N = C^N])]$$

(This is Fermat's last theorem.) ■

EXAMPLE 12.40
Different ways to say things

One of the interesting characteristics of predicate calculus is that there are many ways to say the same thing. For example,

1 Strictly speaking, what we are describing here is the *first-order* predicate calculus. There exist higher-order calculi in which predicates can be applied to predicates, not just to atoms and variables. Prolog allows the user to construct higher-order predicates using `call`; the formalization of such predicates is beyond the scope of this book.

$$\begin{aligned}(P_1 \rightarrow P_2) &\equiv (\neg P_1 \vee P_2)\\ (\neg\exists X[P(X)]) &\equiv (\forall X[\neg P(X)])\\ \neg(P_1 \wedge P_2) &\equiv (\neg P_1 \vee \neg P_2)\end{aligned}$$

This flexibility of expression tends to be handy for human beings, but it can be a nuisance for automatic theorem proving. Propositions are much easier to manipulate algorithmically if they are placed in some sort of *normal form*. One popular candidate is known as *clausal form*. We consider this form in the following section. ■

12.3.1 Clausal Form

As it turns out, clausal form is very closely related to the structure of Prolog programs: once we have a proposition in clausal form, it will be relatively easy to translate it into Prolog. We should note at the outset, however, that the translation is not perfect: there are aspects of predicate calculus that Prolog cannot capture, and there are aspects of Prolog (e.g., its imperative and database-manipulating features) that have no analogues in predicate calculus.

Clocksin and Mellish [CM03, Chap. 10] describe a five-step procedure (based heavily on an article by Martin Davis [Dav63]) to translate an arbitrary first-order predicate proposition into clausal form. We trace that procedure here.

EXAMPLE 12.41
Conversion to clausal form

In the first step, we eliminate implication and equivalence operators. As a concrete example, the proposition

$$\forall A[\neg \mathsf{student}(A) \rightarrow (\neg \mathsf{dorm\_resident}(A) \wedge \neg\exists B[\mathsf{takes}(A, B) \wedge \mathsf{class}(B)])]$$

would become

$$\forall A[\mathsf{student}(A) \vee (\neg \mathsf{dorm\_resident}(A) \wedge \neg\exists B[\mathsf{takes}(A, B) \wedge \mathsf{class}(B)])]$$

In the second step, we move negation inward so that the only negated items are individual terms (predicates applied to arguments):

$$\begin{aligned}&\forall A[\mathsf{student}(A) \vee (\neg \mathsf{dorm\_resident}(A) \wedge \forall B[\neg(\mathsf{takes}(A, B) \wedge \mathsf{class}(B))])]\\ \equiv\ &\forall A[\mathsf{student}(A) \vee (\neg \mathsf{dorm\_resident}(A) \wedge \forall B[\neg \mathsf{takes}(A, B) \vee \neg \mathsf{class}(B)])]\end{aligned}$$

In the third step, we use a technique known as Skolemization (due to logician Thoralf Skolem) to eliminate existential quantifiers. We will consider this technique further in Section C-12.3.3. Our example has no existential quantifiers at this stage, so we proceed.

In the fourth step, we move all universal quantifiers to the outside of the proposition (in the absence of naming conflicts, this does not change the proposition's meaning). We then adopt the convention that all variables are universally quantified, and drop the explicit quantifiers:

$$\text{student}(A) \vee (\neg\text{dorm\_resident}(A) \wedge (\neg\text{takes}(A, B) \vee \neg\text{class}(B)))$$

Finally, in the fifth step, we use the distributive, associative, and commutative rules of Boolean algebra to convert the proposition to *conjunctive normal form*, in which the operators $\wedge$ and $\vee$ are nested no more than two levels deep, with $\wedge$ on the outside and $\vee$ on the inside:

$$(\text{student}(A) \vee \neg\text{dorm\_resident}(A)) \wedge (\text{student}(A) \vee \neg\text{takes}(A, B) \vee \neg\text{class}(B))$$

Our proposition is now in clausal form. Specifically, it is in conjunctive normal form, with negation only of individual terms, with no existential quantifiers, and with implied universal quantifiers for all variables (i.e., for all names that are neither constants nor predicates). The clauses are the items at the outer level: the things that are and-ed together. ■

EXAMPLE 12.42
Conversion to Prolog

To translate the proposition to Prolog, we convert each logical clause to a Prolog fact or rule. Within each clause, we use commutativity to move the negated terms to the right and the non-negated terms to the left (our example is already in this form). We then note that we can recast the disjunctions as implications:

$$\begin{aligned} & (\text{student}(A) \leftarrow \neg(\neg\text{dorm\_resident}(A))) \\ & \qquad \wedge (\text{student}(A) \leftarrow \neg(\neg\text{takes}(A, B) \vee \neg\text{class}(B))) \\ \equiv\ & (\text{student}(A) \leftarrow \text{dorm\_resident}(A)) \\ & \qquad \wedge (\text{student}(A) \leftarrow (\text{takes}(A, B) \wedge \text{class}(B))) \end{aligned}$$

These are Horn clauses. The translation to Prolog is trivial:

```
student(A) :- dorm_resident(A).
student(A) :- takes(A, B), class(B).
```
■

12.3.2 Limitations

We claimed at the beginning of Section 12.1 that Horn clauses could be used to capture most, though not all, of first-order predicate calculus. So what is it missing? What can go wrong in the translation? The answer has to do with the number of non-negated terms in each clause. If a clause has more than one, then if we attempt to cast it as an implication there will be a disjunction on the left-hand side of the $\leftarrow$ symbol, something that isn't allowed in a Horn clause. Similarly, if we end up with no non-negated terms, then the result is a headless Horn clause, something that Prolog allows only as a query, not as an element of the database.

EXAMPLE 12.43
Disjunctive left-hand side

As an example of a disjunctive head, consider the statement "every living thing is an animal or a plant." In clausal form, we can capture this as

$$\text{animal}(X) \vee \text{plant}(X) \vee \neg\text{living}(X)$$

or equivalently

$$\text{animal}(X) \vee \text{plant}(X) \leftarrow \text{living}(X)$$

Because we are restricted to a single term on the left-hand side of a rule, the closest we can come to this in Prolog is

```
animal(X) :- living(X), \+(plant(X)).
plant(X) :- living(X), \+(animal(X)).
```

But this is not the same, because Prolog's \+ indicates inability to prove, not falsehood. ■

EXAMPLE 12.44
Empty left-hand side

As an example of an empty head, consider Fermat's last theorem (Example C-12.39). Abstracting out the math, we might write

$$\forall N[\text{big}(N) \rightarrow \neg(\exists A, \exists B, \exists C[\text{works}(A, B, C, N)])]$$

which becomes the following in clausal form:

$$\neg\text{big}(N) \vee \neg\text{works}(A, B, C, N)$$

We can couch this as a Prolog query:

```
?- big(N), works(A, B, C, N).
```

(a query that will never terminate), but we cannot express it as a fact or a rule. ■

EXAMPLE 12.45
Theorem proving as a search for contradiction

The careful reader may have noticed that facts are entered on the left-hand side of an (implied) Prolog `:-` sign:

```
rainy(rochester).
```

while queries are entered on the right:

```
?- rainy(rochester).
```

The former means

$$\text{rainy(rochester)} \leftarrow \text{true}$$

The latter means

$$\text{false} \leftarrow \text{rainy(rochester)}$$

If we apply resolution to these two propositions, we end up with the contradiction

$$\text{false} \leftarrow \text{true}$$

This observation suggests a mechanism for automated theorem proving: if we are given a collection of axioms and we want to prove a theorem, we temporarily add the *negation* of the theorem to the database and then attempt, through a series of resolution operations, to obtain a contradiction. ■

12.3.3 Skolemization

EXAMPLE 12.46
Skolem constants

In Example C-12.41 we were able to translate a proposition from predicate calculus into clausal form without worrying about existential quantifiers. But what about a statement like this one:

$$\exists X[\mathsf{takes}(X, \mathsf{cs254}) \wedge \mathsf{class\_year}(X, 2)]$$

(There is at least one sophomore in cs254.) To get rid of the existential quantifier, we can introduce a *Skolem constant* x:

$$\mathsf{takes}(\mathsf{x}, \mathsf{cs254}), \mathsf{class\_year}(\mathsf{x}, 2)$$

The mathematical justification for this change is based on something called the *axiom of choice*; intuitively, we say that if there exists an X that makes the statement true, then we can simply pick one, name it x, and proceed. (If there does not exist an X that makes the statement true, then we can choose some arbitrary x, and the statement will still be false.) It is worth noting that Skolem constants are not necessarily distinct; it is quite possible, for example, for x to name the same student as some other constant y that represents a sophomore in `his201`. ■

Sometimes we can replace an existentially quantified variable with an arbitrary constant x. Often, however, we are constrained by some surrounding universal quantifier. Consider the following example:

EXAMPLE 12.47
Skolem functions

$$\forall X[\neg\mathsf{dorm\_resident}(X) \vee \exists A[\mathsf{campus\_address\_of}(X, A)]]$$

(Every dorm resident has a campus address.) To get rid of the existential quantifier, we must choose an address for X. Since we don't know who X is (this is a general statement about all dorm residents), we must choose an address that *depends on X*:

$$\forall X[\neg\mathsf{dorm\_resident}(X) \vee \mathsf{campus\_address\_of}(X, \mathsf{f}(X))]$$

Here f is a *Skolem function*. If we used a simple Skolem constant instead, we'd be saying that there exists some single address shared by all dorm residents. ■

EXAMPLE 12.48
Limitations of Skolemization

Whether Skolemization results in a clausal form that we can translate into Prolog depends on whether we need to know what the constant is. If we are using predicates `takes` and `class_year`, and we wish to assert as a fact that there is a sophomore in `cs254`, we can write

```
takes(the_distinguished_sophomore_in_254, cs254).
class_year(the_distinguished_sophomore_in_254, 2).
```

Similarly, we can assert that every dorm resident has a campus address by writing

```
campus_address_of(X, the_dorm_address_of(X)) :- dorm_resident(X).
```

Now we can search for classes with sophomores in them:

```
sophomore_class(C) :- takes(X, C), class_year(X, 2).
?- sophomore_class(C).
C = cs254
```

and we can search for people with campus addresses:

```
has_campus_address(X) :- campus_address_of(X, Y).
dorm_resident(li_ying).
?- has_campus_address(X).
X = li_ying
```

Unfortunately, we won't be able to identify a sophomore in `cs254` by name, nor will we be able to identify the address of `li_ying`. ■

✓ CHECK YOUR UNDERSTANDING

15. Define the notion of *clausal form* in predicate calculus.
16. Outline the procedure to convert an arbitrary predicate calculus statement into clausal form.
17. Characterize the statements in clausal form that cannot be captured in Prolog.
18. What is *Skolemization*? Explain the difference between Skolem constants and Skolem functions.
19. Under what circumstances may Skolemization fail to produce a clausal form that can be captured usefully in Prolog?

12 Logic Languages

12.6 Exercises

12.19 Restate the following Prolog rule in predicate calculus, using appropriate quantifiers:

```
sibling(X, Y) :- mother(M, X), mother(M, Y),
                 father(F, X), father(F, Y).
```

12.20 Consider the following statement in predicate calculus:

$$\text{empty\_class}(C) \leftarrow \neg\exists X[\text{takes}(X, C)]$$

(a) Translate this statement to clausal form.

(b) Can you translate the statement into Prolog? Does it make a difference whether you're allowed to use \+?

(c) How about the following:

$$\text{takes\_everything}(X) \leftarrow \forall C[\text{takes}(X, C)]$$

Can this be expressed in Prolog?

12.21 Consider the seemingly contradictory statement

$$\neg\text{foo(X)} \rightarrow \text{foo(X)}$$

Convert this statement to clausal form, and then translate into Prolog. Explain what will happen if you ask

```
?- foo(bar).
```

Now consider the straightforward translation, without the intermediate conversion to clausal form:

```
foo(X) :- \+(foo(X)).
```

Now explain what will happen if you ask

```
?- foo(bar).
```

12 Logic Languages

12.7 Explorations

12.27 In Section C-12.3.1 we translated propositions into *conjunctive normal form*: the AND of a collection of ORs. One can also translate propositions into *disjunctive normal form*: the OR of a collection of ANDs. Does disjunctive normal form have any useful properties? What other normal forms exist in mathematical logic? What are their uses?

12.28 With all the different ways to express the same proposition in predicate calculus, is there any useful notion of a "simplest" form? Is it possible, for example, to find, among all equivalent propositions, the one with the smallest number of symbols? How difficult is this task?

12.29 *Satisfiability* is the canonical NP-complete problem. Given a formula in propositional logic (no predicates or quantifiers), it asks whether there exists an assignment of truth values to variables that makes the overall proposition true. Can we use Prolog to solve the satisfiability problem? If not, why not? If so, given that it has to take exponential time, how can we hope to solve problems full of predicates and quantifiers quickly?

12.30 Suppose we had a form of "constructive negation" in Prolog that allowed us to capture information of the form $\forall X[\neg \mathsf{P}(X)]$. What might such a feature look like? What would be its implications for the Prolog search strategy? What portions of predicate calculus (if any) would still be inexpressible?

13 Concurrency

13.5 Message Passing

While shared-memory concurrent programming is common on small-scale multicore and multiprocessor machines, most programs that run on clusters, supercomputers, or geographically distributed machines are currently based on messages. In Sections C-13.5.1 through C-13.5.3 we consider three principal issues in message-based computing: naming, sending, and receiving. In Section C-13.5.4 we look more closely at one particular combination of send and receive semantics, namely remote procedure call. Most of our examples will be drawn from the Ada, Erlang, and Go programming languages, the Java network library, and the MPI library package.

13.5.1 Naming Communication Partners

EXAMPLE 13.53
Naming processes, ports, and entries

To send or receive a message, one must generally specify where to send it to, or where to receive it from: communication partners need names for (or references to) one another. Names may refer directly to a thread or process. Alternatively, they may refer to an *entry* or *port* of a module, or to some sort of *socket* or *channel* abstraction. We illustrate these options in Figure C-13.21. ■

The first naming option—addressing messages to processes—appears in Hoare's original CSP (Communicating Sequential Processes) [Hoa78], an influential proposal for simple communication mechanisms. It also appears in Erlang and in MPI. Each MPI process has a unique `id` (an integer), and each `send` or `receive` operation specifies the `id` of the communication partner. MPI implementations are required to be reentrant; a process can safely be divided into multiple threads, each of which can send or receive messages on the process's behalf.

EXAMPLE 13.54
`entry` calls in Ada

The second naming option—addressing messages to ports—appears in Ada. An Ada *entry call* of the form `t.foo(args)` sends a message to the `entry` named `foo` in task (thread) `t` (`t` may be either a task name or the name of a variable

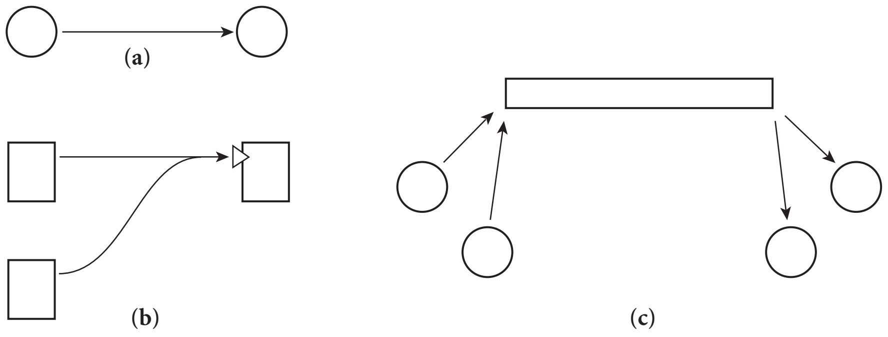


Figure 13.21 Three common schemes to name communication partners. In (a), processes name each other explicitly. In (b), senders name an *input port* of a receiver. The port may be called an *entry* or an *operation*. The receiver is typically a module with one or more threads inside. In (c), senders and receivers both name an independent *channel* abstraction, which may be called a *connection* or a *mailbox*.

whose value is a pointer to a task). As we saw in Section 13.2.3, an Ada task resembles a module; its entries resemble subroutine headers nested directly inside the task. A task receives a message that has been sent to one of its entries by executing an `accept` statement (to be discussed in Section C-13.5.3). Every `entry` belongs to exactly one task; all messages sent to the same `entry` must be received by that one task. ■

EXAMPLE 13.55 Channels in Go

The third naming option—addressing messages to channels—appears in Go and Occam. (Though their concurrency features are loosely based on CSP, both Go and Occam differ from Hoare's proposal in several concrete ways, including the use of channels.) Channel declarations in Go are supported with the `chan` type constructor:

```
var c1 chan int
```

This code declares `c1` to be an (initially `nil`) reference to a channel. A channel value can be created with the built-in function `make`:

```
c1 = make(chan int)
```

Typically the declaration and initialization appear together:

```
var c1 = make(chan int)
```

Here Go infers the type of `c1` from the initialization expression.

To send a message on a channel, a thread uses the binary "arrow" operator `<-` with a channel variable on the left and a message on the right:

```
c1 <- 3
```

To receive, it uses <- as a unary operator, with the channel on the right:

```
my_int = <-c1
```

To indicate that no further messages will be forthcoming, a thread can *close* a channel. A receiving thread can check for this possibility by assigning a receive expression into a pair, the second element of which is a Boolean:

```
my_int, ok = <-c1
if (ok) {
    // use my_int ...
```

■

EXAMPLE 13.56
Remote invocation in Go

For the common idiom in which a server thread is willing to accept requests from any of many possible client threads, each request message can include a reference to the channel on which to send a response:

```
type request struct {
    name string
    reply_to chan string
}
...
// Assume a server thread is listening on chan 'service'
...
var c = make(chan string, 1)    // create channel for response
service <- request{"Alice", c}  // send look-up request for Alice
println(<-c)                    // receive response on c
```

■

Internet Messaging

Java's standard `java.net` library provides two styles of message passing, corresponding to the UDP and TCP Internet protocols. UDP is the simpler of the two. It is a *datagram* protocol, meaning that each message is sent to its destination independently and unreliably. The network software will attempt to deliver it, but makes no guarantees. Moreover two messages sent to the same destination (assuming they both arrive) may arrive in either order. UDP messages use port-based naming (Figure C-13.21b): each message is sent to a specific *Internet address* and *port number*.[1] The TCP protocol also uses port-based naming, but only for the purpose of establishing *connections* (Figure C-13.21c), which it then uses for all subsequent communication. Connections deliver messages reliably and in order.

1 Every publicly visible machine on the Internet has its own unique address. Though a transition to 128-bit addresses has been underway for some time, as of 2008 most addresses are still 32-bit integers, usually printed as four period-separated fields (e.g., 192.5.54.209). Internet name servers translate symbolic names (e.g., `gate.cs.rochester.edu`) into numeric addresses. Port numbers are also integers, but are local to a given Internet address. Ports 1024 through 4999 are generally available for application programs; larger and smaller numbers are reserved for servers.

EXAMPLE 13.57
Datagram messages in Java

To send or receive UDP messages, a Java thread must create a *datagram socket*:

```
DatagramSocket my_socket = new DatagramSocket(port_id);
```

The parameter of the `DatagramSocket` constructor is optional; if it is not specified, the operating system will choose an available port. Typically servers specify a port and clients allow the OS to choose. To send a UDP message, a thread says

```
DatagramPacket my_msg = new DatagramPacket(buf, len, addr, port);
...  // initialize message
my_socket.send(my_msg);
```

The parameters to the `DatagramPacket` constructor specify an array of bytes `buf`, its length `len`, and the Internet address and port of the receiver. Receiving is symmetric:

```
my_socket.receive(my_msg);
...  // parse content of my_msg
```

EXAMPLE 13.58
Connection-based messages in Java

For TCP communication, a server typically "listens" on a port to which clients send requests to establish a connection:

```
ServerSocket my_server_socket = new ServerSocket(port_id);
Socket client_connection = my_server_socket.accept();
```

The `accept` operation blocks until the server receives a connection request from a client. Typically a server will immediately fork a new thread to communicate with the client; the parent thread loops back to wait for another connection with `accept`.

A client sends a connection request by passing the server's symbolic name and port number to the `Socket` constructor:

```
Socket server_connection = new Socket(host_name, port_id);
```

Once a connection has been created, a client and server in Java typically call methods of the `Socket` class to create input and output `streams`, which support all of the standard Java mechanisms for text I/O (Section C-8.7.3):

```
BufferedReader in = new BufferedReader(
    new InputStreamReader(client_connection.getInputStream()));
PrintStream out =
    new PrintStream(client_connection.getOutputStream());
// This is in the server; the client would make streams out
// of server_connection.
...
String s = in.readLine();
out.println("Hi, Mom\n");
```

Among all the message-passing mechanisms we have considered, datagrams are the only one that does not provide some sort of *ordering* constraint. In general, most message-passing systems guarantee that messages sent over the same "communication path" arrive in order. When naming processes explicitly, a path links a single sender to a single receiver. All messages from that sender to that receiver arrive in the order sent. When naming ports, a path links an arbitrary number of senders to a single receiver. Messages that arrive at a port in a given order will be seen by receivers in that order. Note, however, that while messages from the same sender will arrive at a port in order, messages from *different* senders may arrive in arbitrary orders.[2] When naming channels, a path links all the senders that can use the channel to all the receivers that can use it. A Java TCP connection has a single OS process at each end, but there may be many threads inside, each of which can use its process's end of the connection. The connection functions as a queue: send (enqueue) and receive (dequeue) operations are ordered, so that everything is received in the order it was sent.

13.5.2 Sending

One of the most important issues to be addressed when designing a `send` operation is the extent to which it may block the caller: once a thread has initiated a `send` operation, when is it allowed to continue execution? Blocking can serve at least three purposes:

Resource management: A sending thread should not modify outgoing data until the underlying system has copied the old values to a safe location. Most systems block the sender until a point at which it can safely modify its data, without danger of corrupting the outgoing message.

Failure semantics: Particularly when communicating over a long-distance network, message passing is more error-prone than most other aspects of computing. Many systems block a sender until they are able to guarantee that the message will be delivered without error.

Return parameters: In many cases a message constitutes a *request*, for which a *reply* is expected. Many systems block a sender until a reply has been received.

When deciding how long to block, we must consider synchronization semantics, buffering requirements, and the reporting of run-time errors.

2 Suppose, for example, that process *A* sends a message to port *p* of process *B*, and then sends a message to process *C*, while process *C* first receives the message from *A* and then sends its own message to port *p* of *B*. If messages are sent over a network with internal delays, and if *A* is allowed to send its message to *C* before its first message has reached port *p*, then it is possible for *B* to hear from *C* before it hears from *A*. This apparent reversal of ordering could easily happen on the Internet, for example, if the message from *A* to *B* traverses a satellite link, while the messages from *A* to *C* and from *C* to *B* use ocean-floor fiber-optic cables.

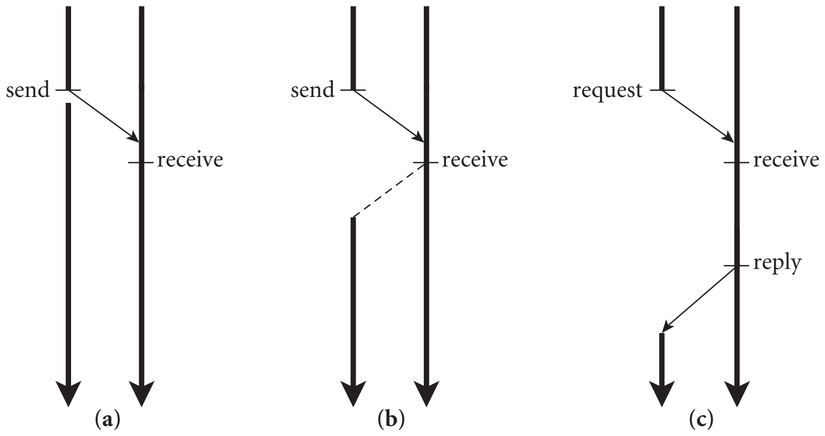


Figure 13.22 **Synchronization semantics for the `send` operation: no-wait `send` (a), synchronization `send` (b), and remote-invocation `send` (c).** In each diagram we have assumed that the original message arrives before the receiver executes its `receive` operation; this need not in general be the case.

Synchronization Semantics

On its way from a sender to a receiver, a message may pass through many intermediate steps, particularly if traversing the Internet. It first descends through several layers of software on the sender's machine, then through a potentially large number of intermediate machines, and finally up through several layers of software on the receiver's machine. We could imagine unblocking the sender after any of these steps, but most of the options would be indistinguishable in terms of user-level program behavior. If we assume for the moment that a message-passing system can always find buffer space to hold an outgoing message, then our three rationales for delay suggest three principal semantic options:

EXAMPLE 13.59
Three main options for `send` semantics

No-wait send: The sender does not block for more than a small, bounded period of time. The message-passing implementation copies the message to a safe location and takes responsibility for its delivery.

Synchronization send: The sender waits until its message has been received.

Remote-invocation send: The sender waits until it receives a reply.

These three alternatives are illustrated in Figure C-13.22. ■

No-wait `send` appears in Erlang and in the Java Internet library. Synchronization `send` appears in Occam and, by default, in Go. (If a Go channel is declared with an explicit *buffering capacity*, however, no-wait `send` is used.) Remote-invocation `send` appears in Ada and in Occam. MPI provides an implementation-oriented hybrid of no-wait `send` and synchronization `send`: a `send` operation blocks until the data in the outgoing message can safely be modified. In implementations that do their own internal buffering, this rule amounts to no-wait `send`. In other implementations, it amounts to synchronization `send`.

The programmer has the option, if desired, to insist on no-wait `send` or synchronization `send`; performance may suffer on some systems if the request is different from the default.

Buffering

In practice, unfortunately, no message-passing system can provide a version of `send` that never waits (unless of course it simply throws some messages away). If we imagine a thread that sits in a loop sending messages to a thread that never receives them, we quickly see that unlimited amounts of buffer space would be required. At some point, any implementation must be prepared to block an overactive sender, to keep it from overwhelming the system. Such blocking is a form of *backpressure*. Milder backpressure can also be applied by reducing a thread's scheduling priority or by increasing the (still bounded) delay before a "no-wait" `send` returns.

EXAMPLE 13.60
Buffering-dependent deadlock

For any fixed amount of buffer space, it is possible to design a program that requires a larger amount of space to run correctly. Imagine, for example, that the message-passing system is able to buffer n messages on a given communication path. Now imagine a program in which A sends $n+1$ messages to B, followed by one message to C. C then sends one message to B, on a different communication path. Finally, B insists on receiving the message from C before receiving the messages from A. If A blocks after message n, implementation-dependent deadlock will result. The best that an implementation can do is to provide a sufficiently

DESIGN & IMPLEMENTATION

13.10 The semantic impact of implementation issues

The inability to buffer unlimited amounts of data and, likewise, to report errors synchronously to a sender that has continued execution are only the most recent of many examples we have seen in which pragmatic implementation issues may restrict the language semantics available to the programmer. Other examples include limitations on the length of source lines or variable names (Section 2.1.1); limits on the memory available for data (whether global, stack, or heap allocated) and for recursive function evaluation (Section 3.2); the lack of ranges in `case` statement labels (Section 6.4.2); `in reverse`, `downto`, and constant step sizes for `for` loops (Section 6.5.1); limits on set universe size (to accommodate bit vectors—Section 8.4); limited procedure nesting (to accommodate displays—Section 9.1); the pointer-only restriction on opaque exports in Modula-2 (Section 10.2.1); and the lack of nested threads or of unrestricted arms on a `cobegin` statement (to avoid the need for cactus stacks—Section 9.5.1). Some of these limitations are reflected in the formal semantics of the language. Others (generally those that vary most from one implementation to another) restrict the set of semantically valid programs that the system will run correctly.

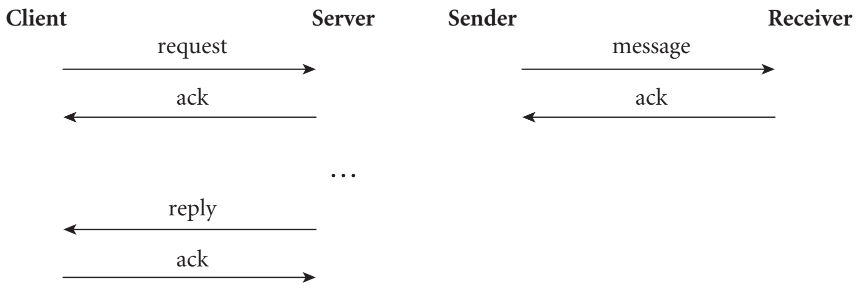


Figure 13.23 Acknowledgment messages for error detection. In the absence of piggybacking, remote-invocation `send` (left) may require four underlying messages; synchronization `send` (right) may require two.

large amount of space that realistic applications are unlikely to find the limit to be a problem. ■

For synchronization `send` and remote-invocation `send`, buffer space is not generally a problem: the total amount of space required for messages is bounded by the number of threads, and there are already likely to be limits on how many threads a program can create. A thread that sends a reply message can always be permitted to proceed: we know that we shall be able to reuse the buffer space quickly, because the thread that sent the request is already waiting for the reply.

Error Reporting

EXAMPLE 13.61
Acknowledgments

If the underlying message-passing system is unreliable, a language or library will typically employ *acknowledgment* messages to verify successful transmission (Figure C-13.23). If an acknowledgment is not received within a reasonable amount of time, the implementation will typically resend. If several attempts fail to elicit an acknowledgment, an error will be reported. ■

As long as the sender of a message is blocked, errors that occur in attempting to deliver a message can be reflected back as exceptions, or as status information in result parameters or global variables. Once a sender has continued, there is no obvious way in which to report any problems that arise. Like limits on message buffering, this dilemma poses semantic problems for no-wait `send`. For UDP, the solution is to state that messages are unreliable: if something goes wrong, the message is simply lost, silently. For TCP, the "solution" is to state that only "catastrophic" errors will cause a message to be lost, in which case the connection will become unusable and future calls will fail immediately. An even more drastic approach was taken in the original version of MPI: certain implementation-specific errors could be detected and handled at run time, but in general if a message could not be delivered then the program as a whole was considered to have failed. Newer versions of MPI provide a richer set of error-reporting facilities that can be used, with some effort, to build fault-tolerant programs.

Emulation of Alternatives

All three varieties of `send` can be emulated by the others. To obtain the effect of remote-invocation `send`, a thread can follow a no-wait `send` of a request with a `receive` of the reply, as we saw in Example C-13.56. Similar code will allow us to emulate remote-invocation `send` using synchronization `send`. To obtain the effect of synchronization `send`, a thread can follow a no-wait `send` with a `receive` of a high-level acknowledgment, which the receiver will send immediately upon receipt of the original message. To obtain the effect of synchronization `send` using remote-invocation `send`, a thread that receives a request can simply reply immediately, with no return parameters.

To obtain the effect of no-wait `send` using synchronization `send` or remote-invocation `send`, we must interpose a buffer process (the message-passing analogue of our shared-memory bounded buffer) that replies immediately to “senders” or “receivers” whenever possible. The space available in the buffer process makes explicit the resource limitations that are always present below the surface in implementations of no-wait `send`.

Syntax and Language Integration

In the emulation examples above, our hypothetical syntax assumed a library-based implementation of message passing. Because `send`, `receive`, `accept`, and so on are ordinary subroutines in such an implementation, they usually take a

DESIGN & IMPLEMENTATION

13.11 Emulation and efficiency

Unfortunately, user-level emulations of alternative `send` semantics are seldom as efficient as optimized implementations using the underlying primitives. Suppose for example that we wish to use remote-invocation `send` to emulate synchronization `send`. Suppose further that our implementation of remote-invocation `send` is built on top of network software that needs acknowledgments to verify message delivery. After sending a reply, the server’s run-time system will wait for an acknowledgment from the client. If a server thread can work for an arbitrary amount of time before sending a reply, then the run-time system will need to send separate acknowledgments for the request and the reply. If a programmer uses this implementation of remote-invocation `send` to emulate synchronization `send`, then the underlying network may end up transmitting a total of four messages (more if there are any transmission errors). By contrast, a “native” implementation of synchronization `send` would require only two underlying messages. In some cases the run-time system for remote-invocation `send` may be able to delay transmission of the first acknowledgment long enough to “piggy-back” it on the subsequent reply if there is one; in this case an emulation of synchronization `send` may transmit three underlying messages instead of only two. We consider the efficiency of emulations further in Exercise C-13.36 and Exploration C-13.52.

fixed, static number of parameters, two of which typically specify the location and size of the message to be sent. To send a message containing values held in more than one program variable, the programmer may need to explicitly *gather*, or *marshal*, those values into the fields of a record. On the receiving end, the programmer may then need to *scatter* (*unmarshal*) the values back into program variables. By contrast, a concurrent programming language can provide message-passing operations whose "argument" lists can include an arbitrary number of values to be sent. Moreover, the compiler can arrange to perform type checking on those values, using techniques similar to those employed for subroutine linkage across compilation units (to be described in Section 15.6.2). Finally, as we shall see in Section C-13.5.3, an explicitly concurrent language can employ non-procedure-call syntax, for example to couple a remote-invocation `accept` and `reply` in such a way that the `reply` doesn't have to explicitly identify the `accept` to which it corresponds.

13.5.3 Receiving

Probably the most important dimension on which to categorize mechanisms for receiving messages is the distinction between explicit `receive` operations and the *implicit* receipt described in Section 13.2.3. Among the languages and systems we have been using as examples, none provides implicit receipt, but it appears in a variety of research languages, and in some of the RPC systems we will consider in Section C-13.5.4).

With implicit receipt, every message that arrives at a given port (or over a given channel) will create a new thread of control, subject to resource limitations (any implementation will have to stall incoming requests when the number of threads grows too large). With explicit receipt, a message will be queued until some already-existing thread indicates a willingness to receive it. At any given point in time there may be a potentially large number of messages waiting to be received. Most languages and libraries with explicit receipt allow a thread to exercise some sort of *selectivity* with respect to which messages it wants to consider.

In MPI, every message includes the `id` of the process that sent it, together with an integer *tag* specified by the sender. A `receive` operation specifies a desired sender `id` and message tag. Only matching messages will be received. In many cases receivers specify "wild cards" for the sender `id` and/or message tag, allowing any of a variety of messages to be received. Special versions of `receive` also allow a process to test (without blocking) to see if a message of a particular type is currently available (this operation is known as *polling*), or to "time out" and continue if a matching message cannot be received within a specified interval of time.

Because they are languages instead of library packages, Ada, Erlang, Go, and Occam are able to use special, non-procedure-call syntax for selective message receipt. Moreover because messages are built into the naming and typing system, these languages are able to receive selectively on the basis of port/channel names

```
task buffer is
    entry insert(d : in bdata);
    entry remove(d : out bdata);
end buffer;

task body buffer is
    SIZE : constant integer := 10;
    subtype index is integer range 1..SIZE;
    buf : array (index) of bdata;
    next_empty, next_full : index := 1;
    full_slots : integer range 0..SIZE := 0;
begin
    loop
        select
          when full_slots < SIZE =>
            accept insert(d : in bdata) do
                buf(next_empty) := d;
            end;
            next_empty := next_empty mod SIZE + 1;
            full_slots := full_slots + 1;
        or
          when full_slots > 0 =>
            accept remove(d : out bdata) do
                d := buf(next_full);
            end;
            next_full := next_full mod SIZE + 1;
            full_slots := full_slots - 1;
        end select;
    end loop;
end buffer;
```

Figure 13.24 Bounded buffer in Ada, with an explicit manager task.

and parameters, rather than the more primitive notion of tags. In all four languages, the selective `receive` construct is a special form of *guarded command*, as described in Section C-6.7.

EXAMPLE 13.62
Bounded buffer in Ada 83

Figure C-13.24 contains code for a bounded buffer in Ada 83. Here an active "manager" thread executes a `select` statement inside a loop. (Recall that it is also possible to write a bounded buffer in Ada using *protected objects*, without a manager thread, as described in Section 13.4.2.) The Ada `accept` statement receives the `in` and `in out` parameters (Section 9.3.1) of a remote invocation request. At the matching `end`, `accept` returns the `in out` and `out` parameters as a reply message. A client task would communicate with the bounded buffer using an *entry call*:

```
-- producer:                    -- consumer:
buffer.insert(3);               buffer.remove(x);
```

The `select` statement in our buffer example has two arms. The first arm may be selected when the buffer is not full and there is an available `insert` request; the second arm may be selected when the buffer is not empty and there is an available `remove` request. Selection among arms is a two-step process: first the guards (`when` expressions) are evaluated, then for any that are true the subsequent `accept` statements are considered to see if a message is available. (The guard in front of an `accept` is optional; if missing it behaves as `when true =>`.) If both of the guards in our example are true (the buffer is partly full) and both kinds of messages are available, then either arm of the statement may be executed, at the discretion of the implementation. (For a discussion of issues of *fairness* in the choice among true guards, see Sidebar C-6.11.) ■

EXAMPLE 13.63
Timeout and distributed termination

Every `select` statement must have at least one arm beginning with `accept` (and optionally `when`). In addition, it may have three other types of arms:

```
when condition => delay how_long
    other_statements
...
or when condition => terminate
...
else ...
```

A `delay` arm may be selected if no other arm becomes selectable within *how_long* seconds. (Ada implementations are required to support delays as long as 1 day or as short as 20 ms.) A `terminate` arm may be selected only if all potential communication partners have already terminated or are likewise stuck in `select` statements with `terminate` arms. Selection of the arm causes the task that was executing the `select` statement to terminate. An `else` arm, if present, will be selected when none of the guards are true or when no `accept` statement can be executed immediately. A `select` statement with an `else` arm is not permitted to have any `delay` arms. In practice, one would probably want to include a `terminate` arm in the `select` statement of a manager-style bounded buffer. ■

EXAMPLE 13.64
Bounded buffer in Go

In Go, a bounded buffer is trivial: it's just a buffered channel:

```
type bdata struct {
    n int   // or whatever
}
var buffer = make(chan bdata, 10)   // space for ten items of type bdata
...
buffer <- bdata{3}                  // insert
...
my_int = (<-buffer).n               // remove
```

To illustrate language features, we can also build a bounded buffer with an explicit thread, an array, and a pair of default (unbuffered) channels, in a manner similar to the Ada example of Figure C-13.24, but with synchronization `send` instead of remote invocation. Code for this alternative appears in Figure C-13.25.

Unlike built-in buffered channels, it could easily be augmented to support functionality like priority-based (as opposed to FIFO) queueing, or methods to clear the buffer or to query the number of messages currently queued. To use the basic `insert/remove` operations, we might write:

```
var b = make_buffer()
...
b.insert(bdata{3})                  // insert
...
my_int = b.remove().n               // remove
```

As in the Ada example, requests are processed by an active manager thread (called a "goroutine" in Go), here started with the `go` command. The `select` statement in Go does not support explicit guards; we have achieved a similar effect in Figure C-13.25 by setting the `ic` and `rc` channels to `nil` when they should not be selected. Because we have used synchronization `send`—channels `insert_c` and `remove_c` have zero capacity—there is an asymmetry between the handling of `insert` and `remove` requests: the former need only send the manager data; the latter must send a channel reference and then wait for the manager to send the data back. ■

EXAMPLE 13.65
Bounded buffer in Erlang

In Erlang, which uses no-wait `send`, one might at first expect asymmetry similar to that of Figure C-13.25: a consumer would have to receive a reply from a bounded buffer, but a producer could simply send data. Such asymmetry would have a hidden flaw, however: because a process does not wait after sending, the producer could easily send more items than the buffer can hold, with the excess being buffered in the message system. If we want the buffer to truly be bounded, we must require the producer to wait for an acknowledgment. Code for the buffer appears in Figure C-13.26. Because Erlang is a functional language, we use tail recursion instead of iteration. Code for the producer and consumer looks like this:

```
-- producer:                      -- consumer:
Buffer ! {insert, X, self()},     Buffer ! {remove, self()},
receive ok -> [] end.             receive X -> [] end.
```

The exclamation point (`!`), borrowed from CSP, is used to send a message. ■

EXAMPLE 13.66
Peeking at messages in Erlang

Several languages—Erlang among them—place the parameters of an incoming message within the scope of the guard condition, allowing a receiver to "peek inside" a message before deciding whether to receive it. In Erlang, we can say

```
receive
    {insert, D} when D rem 2 == 1 ->    % accept only odd numbers
```

The ability to peek implies that the content of incoming messages must be visible to the language run-time system. An Erlang implementation must therefore be prepared to accept (and buffer) an arbitrary number of messages; it cannot rely on the operating system or other underlying software to provide the buffering for it. Moreover the fact that buffer space can never be truly unlimited means that guards and scheduling expressions will be unable to see messages whose delivery has been delayed by backpressure. ■

```
type buffer struct {
    full_slots, next_full, next_empty int
    buf [SIZE]bdata
    insert_c chan bdata
    remove_c chan chan bdata
}
func manager(b *buffer) {
    var ic chan bdata = b.insert_c
    var rc chan chan bdata = nil
    for {           // at least one of ic and rc will always be non-nil
        select {
          case d := <-ic:       // := means "declare and initialize"
            b.buf[b.next_empty] = d
            b.next_empty = (b.next_empty + 1) % SIZE
            b.full_slots++
            rc = b.remove_c     // there is definitely data to remove
            if b.full_slots == SIZE { ic = nil }
          case c := <-rc:
            c <- b.buf[b.next_full]
            b.next_full = (b.next_full + 1) % SIZE
            b.full_slots--
            ic = b.insert_c     // there is definitely space to fill
            if b.full_slots == 0 { rc = nil }
        }
    }
}
func make_buffer() (b *buffer) {    // return value has name 'b'
    b = new(buffer)
    b.full_slots = 0
    b.next_full = 0
    b.next_empty = 0
    b.insert_c = make(chan bdata)
    b.remove_c = make(chan chan bdata)
    go manager(b)               // create active manager thread
    return
}
func (b *buffer) insert(e bdata) {
    b.insert_c <- e             // send data to manager
}
func (b *buffer) remove() bdata {
    var c = make(chan bdata)
    b.remove_c <- c             // send temporary channel to manager
    return <-c                  // receive and return response
}
```

Figure 13.25 Bounded buffer with an explicit manager thread in Go. The `insert` and `remove` functions serve as methods of buffer `b`. Note that in the absence of additional functionality (not shown), this code would better be replaced by trivial use of a buffered channel with capacity `SIZE`. Also, if using this version, we would probably want a way to terminate the manager thread when the buffer is no longer needed.

```
buffer(Max, Free, Q) ->
    receive
        {insert, D, Client} when Free > 0 ->
            Client ! ok,                             % send ack
            buffer(Max, Free-1, queue:in(D, Q));     % enqueue
        {remove, Client} when Free < Max ->
            {{value, D}, NewQ} = queue:out(Q),       % dequeue
            Client ! D,                              % send element
            buffer(Max, Free+1, NewQ)
    end.
```

Figure 13.26 **Bounded buffer in Erlang.** Variables (names that can be instantiated with a value) begin with a capital letter; constants begin with a lower-case letter. Queue operations (`in`, `out`) are provided by the standard Erlang library. Typing is dynamic. The `send` operator (`!`) is as in CSP and Occam. Each clause of the `receive` ends with a tail recursive call.

13.5.4 Remote Procedure Call

Any of the three principal forms of `send` (no-wait, synchronization, remote-invocation) can be paired with either of the principal forms of `receive` (explicit or implicit). The combination of remote-invocation `send` with explicit receipt (e.g., as in Ada) is sometimes known as *rendezvous*. The combination of remote-invocation `send` with implicit receipt is usually known as *remote procedure call*. RPC is available in several concurrent languages, and is also supported on many systems by augmenting a sequential language with a *stub compiler*. The stub compiler is independent of the language's regular compiler. It accepts as input a formal description of the subroutines that are to be called remotely. The description is roughly equivalent to the subroutine headers and declarations of the types of all parameters. Based on this input the stub compiler generates source code for *client* and *server stubs*. A client stub for a given subroutine marshals request parameters and an indication of the desired operation into a message buffer, sends the message to the server, waits for a reply message, and unmarshals that message into result parameters. A server stub takes a message buffer as parameter, unmarshals request parameters, calls the appropriate local subroutine, marshals return parameters into a reply message, and sends that message back to the appropriate client. Invocation of a client stub is relatively straightforward. Invocation of server stubs is discussed under "Implementation" below.

Semantics

A principal goal of most RPC systems is to make the remote nature of calls as *transparent* as possible; that is, to make remote calls look as much like local calls as possible [BN84]. In a stub compiler system, a client stub should have the same interface as the remote procedure for which it acts as proxy; the programmer should usually be able to call the routine without knowing or caring whether it is local or remote.

Several issues make it difficult to achieve transparency in practice:

Parameter modes: It is difficult to implement call-by-reference parameters across a network, since actual parameters will not be in the address space of the called routine. (Access to global variables is similarly difficult.)

Performance: There is no escaping the fact that remote procedures may take a long time to return. In the face of network delays, one cannot use them casually.

Failure semantics: Remote procedures are much more likely to fail than are local procedures. It is generally acceptable in the local case to assume that a called procedure will either run exactly once or else the entire program will fail. Such an assumption is overly restrictive in the remote case.

We can use value/result parameters in place of reference parameters so long as program correctness does not rely on the aliasing created by reference parameters. As noted in Section 9.3.1, Ada declares that a program is *erroneous* if it can tell the difference between pass-by-reference and pass-by-value/result implementations of `in out` parameters. If absolutely necessary, reference parameters and global variables can be implemented with message-passing thunks in a manner reminiscent of call-by-name parameters (Section C-9.3.2), but only at very high cost. As noted in Section 7.4, a few languages and systems perform deep copies of linked data structures passed to remote routines.

Performance differences between local and remote calls can be hidden only by artificially slowing down the local case. Such an option is clearly unacceptable.

DESIGN & IMPLEMENTATION

13.12 Parameters to remote procedures

Ada's comparatively high-level semantics for parameter modes allows the same set of modes to be used for both subroutines and entries (rendezvous). An Ada compiler will generally pass a large argument to a subroutine by reference whenever possible, to avoid the expense of copying. If tasks are on separate nodes of a cluster, however, the compiler will generally pass the same argument to an entry by value-result.

A few concurrent languages provide parameter modes specifically designed with remote invocation in mind. In Emerald [BHJL07], for example, every parameter is a reference to an object. References to remote objects are implemented transparently via message passing. To minimize the frequency of such references, objects passed to remote procedures often *migrate* with the call: they are packaged with the request message, sent to the remote site (where they can be accessed locally), and returned to the caller in the reply. Emerald calls this *call by move*. In Hermes [SBG$^+$91] and Rust, parameter passing is *destructive*: arguments become uninitialized from the caller's point of view, and can therefore migrate to a remote callee without danger of inducing remote references.

Exactly-once failure semantics can be provided by aborting the caller in the event of failure or, in highly reliable systems, by delaying the caller until the operating system or language run-time system is able to rebuild the failed computation using information previously dumped to disk. (Failure recovery techniques are beyond the scope of this text.) An attractive alternative is to accept "at-most-once" semantics with notification of failure. The implementation retransmits requests for remote invocations as necessary in an attempt to recover from lost messages. It guarantees that retransmissions will never cause an invocation to happen more than once, but it admits that in the presence of communication failures the invocation may not happen at all. If the programming language provides exceptions then the implementation can use them to make communication failures look like any other kind of run-time error.

Implementation

At the level of the kernel interface, `receive` is usually an explicit operation. To make `receive` appear implicit to the application programmer, the code produced by an RPC stub compiler (or the run-time system of an RPC-based language) must bridge this explicit-to-implicit gap. The typical implementation resembles the thread-based event handling of Section 9.6.2. We describe it here in terms of stub compilers; in a concurrent language with implicit receipt the regular compiler does essentially the same work.

EXAMPLE 13.67
An RPC server system

Figure C-13.27 illustrates the layers of a typical RPC system. Code above the upper horizontal line is written by the application programmer. Code in the middle is a combination of library routines and code produced by the RPC stub compiler. To initialize the RPC system, the application makes a pair of calls into the run-time system. The first provides the system with pointers to the stub routines produced by the stub compiler; the second starts a *message dispatcher*. What happens after this second call depends on whether the server is concurrent and, if so, whether its threads are implemented on top of one OS process or several.

In the simplest case—a single-threaded server on a single OS process—the dispatcher runs a loop that calls into the kernel to receive a message. When a message arrives, the dispatcher calls the appropriate RPC stub, which unmarshals request parameters and calls the appropriate application-level procedure. When that procedure returns, the stub marshals return parameters into a reply message, calls into the kernel to send the message back to the caller, and then returns to the dispatcher. ■

This simple organization works well so long as each remote request can be handled quickly, without ever needing to block. If remote requests must sometimes wait for user-level synchronization, then the server's process must manage a ready list of threads, as described in Section 13.2.4, but with the dispatcher integrated into the usual thread scheduler. When the current thread blocks (in application code), the scheduler/dispatcher will grab a new thread from the ready list. If the ready list is empty, the scheduler/dispatcher will call into the kernel to receive a message, fork a new user-level thread to handle it, and then continue to execute

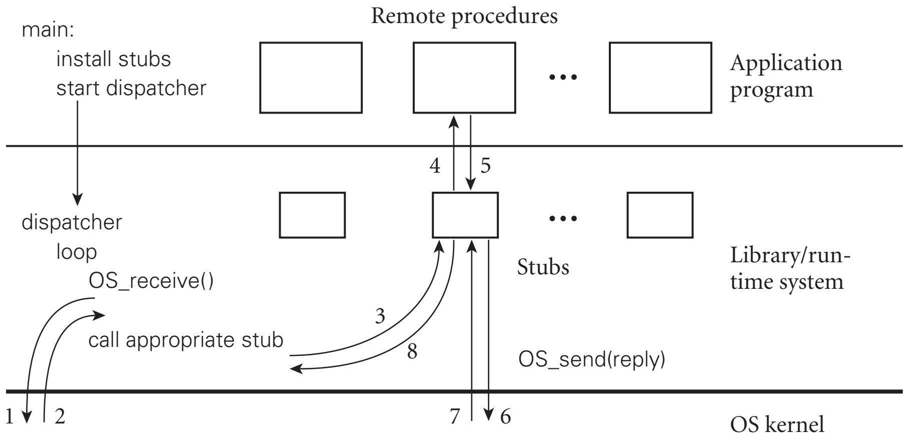


Figure 13.27 Implementation of a remote procedure call server. Application code initializes the RPC system by installing stubs generated by the stub compiler (not shown). It then calls into the run-time system to enable incoming calls. Depending on details of the particular system in use, the dispatcher may use the thread from the main program (in which case the call to start the dispatcher never returns), or it may create a pool of threads that handle incoming requests.

runnable threads until the list is empty again (each thread will terminate when it finishes handling its request).

In a multithreaded server, the call to start the dispatcher will generally ask the kernel to fork a "pool" of threads to service remote requests. Each of these threads will then perform the operations described in the previous paragraphs. In a language or library with a one–one correspondence between user threads and kernel threads, each will repeatedly receive a message from the kernel, call the appropriate stub, and loop back for another request. With a more general thread package, each kernel thread will run threads from the application's ready list until the list is empty, at which point it (the kernel thread) will call into the kernel for another message. So long as the number of runnable user threads is greater than or equal to the number of kernel threads, no new messages will be received. When the number of runnable user threads drops below the number of kernel threads, the extra kernel threads will call into the kernel, where they will block until requests arrive.

✓ CHECK YOUR UNDERSTANDING

50. Describe three ways in which processes or threads commonly name their communication partners.

51. What is a *datagram*?

52. Why, in general, might a `send` operation need to block?
53. What are the three principal synchronization options for the sender of a message? What are the tradeoffs among them?
54. What are *gather* and *scatter* operations in a message-passing program? What are *marshalling* and *unmarshalling*?
55. Describe the tradeoffs between *explicit* and *implicit* message receipt.
56. What is a *remote procedure call* (RPC)? What is a *stub compiler*?
57. What are the obstacles to *transparency* in an RPC system?
58. What is a *rendezvous*? How does it differ from a remote procedure call?
59. Explain the purpose of a `select` statement in Ada or Go.
60. What semantic and pragmatic challenges are introduced by the ability to "peek" inside messages before they are received?

13 Concurrency

13.7 Exercises

13.34 In Section 13.4.1 we cast monitors as a mechanism for synchronizing access to shared memory, and we described their implementation in terms of semaphores. It is also possible to think of a monitor as a module inhabited by a single thread, which accepts request messages from other threads, performs appropriate operations, and replies. Give the details of a monitor implementation consistent with this conceptual model. Be sure to include condition variables. (Hint: See the discussion of early reply in Section 13.2.3.)

13.35 Show how shared memory can be used to implement message passing. Specifically, choose a set of message-passing operations (e.g., no-wait `send` and explicit message receipt) and show how to implement them in your favorite shared-memory notation.

13.36 When implementing reliable messages on top of unreliable messages, a sender can wait for an acknowledgment message, and retransmit if it doesn't receive it within a bounded period of time. But how does the receiver know that its acknowledgment has been received? Why doesn't the sender have to acknowledge the acknowledgment (and the receiver acknowledge the acknowledgment of the acknowledgment ...)? (For more information on the design of fast, reliable protocols, you might want to consult a text on computer networks [TW12, PD12].)

13.37 While Go allows both *input* (receive) and *output* (send) guards on its `select` statements, Occam and CSP allow only input guards. The difference has to do with the fact that Go is designed for communication among threads in a single address space, while Occam and CSP were designed for a distributed environment. Why should this make a difference? Suppose you wished to add output guards to Occam. How would the implementation work? (Hint: For ideas, see the article by Bagrodia [Bag89].)

13.38 In Section C-13.5.3 we described the semantics of a `terminate` arm on an Ada `select` statement: this arm may be selected if and only if all potential communication partners have terminated, or are likewise stuck in `select` statements with `terminate` arms. Erlang and Occam have no similar facility, though the original CSP proposal does. How would you implement `terminate` arms in Ada? Why do you suppose they were left out of Erlang and Occam? (Hint: For ideas, see the work of Apt and Francez [Fra80, AF84].)

13 Concurrency

13.8 Explorations

13.52 Find out how message passing is implemented in some locally available concurrent language or library. Does this system provide no-wait `send`, synchronization `send`, remote-invocation `send`, or some related hybrid? If you wanted to emulate the other options using the one available, how expensive would emulation be, in terms of low-level operations performed by the underlying system? How would this overhead compare to what could be achieved on the same underlying system by a language or library that provided an optimized implementation of the other varieties of `send`?

13.53 MPI provides extensive facilities for *collective communication*, in which there are more than two communicating parties. Examples include *multicast*, in which a message is sent simultaneously to a group of recipients; *scatter*, in which elements of an array-structured message are sent, one each, to a group of recipients; *gather*, in which an array-structured message is created, at the sole recipient, from elements provided by a group of senders; *all-to-all*, in which participants provide one element each of an array-structured message that is received by all; and *reduction*, in which messages from a group of senders are combined, using a commutative operator, into a result that is received by one or all. Learn more about both the semantics and the implementation of collective communication. What opportunities does it provide for optimizations that are difficult to implement at the application level?

13.54 Language designers and concurrency experts have argued for nearly 40 years over whether shared memory or message passing is a more appealing programming model. The argument is to a large extent subjective—and hence not subject to definitive settlement—but it includes substantive issues of fault containment, implementation efficiency, hardware requirements, and algorithmic expressiveness as well. Do a literature search on "shared memory versus message passing." How many papers do you find? Read a sampling of these and summarize their arguments. Do you find any of the positions particularly convincing?

14 Scripting Languages

14.3.5 XSLT

HTML was inspired by an older and more complicated standard known as SGML (standard generalized markup language). SGML was developed in the 1980s by a consortium of government agencies and major corporations, to represent structured data. It was used, for example, to computerize both the Oxford English Dictionary and the technical documentation of Boeing Corp.

In the early days of the Web, SGML was clearly too complex and formal for web pages, which needed to be written by hand and rendered in real time by slow computers. The more informal replacement evolved in an ad hoc way, with the result that HTML has been very difficult to standardize. Incompatibilities among browsers continue to frustrate web designers, and several features of the language that have been deprecated in the most recent standards are nonetheless still widely used. Other features, while not deprecated, are widely regarded in hindsight to have been mistakes.

EXAMPLE 14.81
Content versus presentation in HTML

Probably the biggest problem with HTML is that it does not adequately distinguish between the *content* and the *presentation* (appearance) of a document. As a trivial example, web designers sometimes use `<i>` ... `</i>` tags to request that text be set in an italic font, when `<em>` ... `</em>` (emphasis) would be more appropriate. A browser for the visually impaired might choose to emphasize text with something other than italics, and might render book titles (also often specified with `<i>` ... `</i>`) in some entirely different fashion. More significantly, many web designers use tables (`<table>` ... `</table>`) to control the relative positioning of elements on a page, when the content isn't tabular at all. As the Web extends across cell phones, televisions, tablets, watches, and audio-only devices, the need to distinguish between content and presentation has become essential. ■

This is where XML steps in. A streamlined descendant of SGML, developed by the World Wide Web Consortium in the mid to late 1990s, XML has at least three important advantages over HTML: (1) its syntax and semantics are more regular and consistent, and more consistently implemented across platforms; (2) it is *extensible*, meaning that users can define new tags; (3) it specifies content only, leaving presentation to a companion standard known as XSL (extensible stylesheet language). As noted in the main text, XSLT is a portion of XSL devoted to *transforming* XML: selecting, reorganizing, and modifying tags and the elements they delimit—in effect, scripting the processing of data represented in XML.

Internet Alphabet Soup

Learning about web standards can be a daunting task: there is an enormous number of buzzwords, standards, and multiletter abbreviations. The standards—and the relationships among them—are also moving targets, promulgated by groups whose interests are not always in sync. To start, it may help to note that each of the major markup languages—SGML, HTML, and XML—has a corresponding *stylesheet language*: DSSSL, CSS, and XSL, respectively. A stylesheet language is used to control the presentation of a document, separate from its content. Stylesheet languages are essential for SGML and XML; without them there is no way to know whether a `<RECORD>` represents a database entry, an antique phonograph album, or an Olympic achievement, much less how to display it. HTML is less dependent on stylesheets, but most professionally maintained web sites use CSS to create a uniform "look and feel" across a collection of pages without embedding redundant information in every page.

SGML is still used for large-scale projects in the business world, though many newer projects have chosen to use the simpler XML. HTML continues to evolve (see sidebar C-14.14). HTML5, finalized by the World Wide Wide Consortium in 2014, adds extensive new support for multimedia content, and specifies both general and XML-compliant versions of the syntax.

XML and XHTML

As a general rule, the syntax of XML is simpler than that of SGML or HTML. To allow XML tools (XSLT in particular) to be used to process web pages, the HTML5 standard defines a restricted version of the HTML syntax, known as XHTML. With a few minor exceptions, any web page that can be specified in HTML can also be specified in XHTML, and vice versa. The `content-type` header that precedes a web page when transmitted over the Internet tells the browser which parser to use: `text/html` means "regular" HTML; `application/xhtml+xml` means XHTML. In practice, the principal differences between the notations are that XHTML is harder for human beings to write, because the rules are stricter, and XML parsers are designed to reject (and decline to render) any page that is not *well formed* (syntactically correct). HTML parsers are designed to tolerate—and do something reasonable with—even the worst "tag soup." With some care, it is possible to write pages that will be processed correctly by both HTML and XHTML parsers; such pages are said to use *polyglot markup* (syntax).

In any well-formed XML document (including those written in XHTML), tags must either constitute properly nested, matched pairs, or be explicit singletons, which end with a "`/>`" delimiter. Similarly, the values of *attributes* (key-value pairs embedded within tags) must always be specified with quotes. The following fragment, for example, is well formed (though incomplete) XHTML:

EXAMPLE 14.82
Well-formed XHTML

```
<em><q id="favorite">I defy the tyranny of precedent</q></em><br />
(Clara Barton)
```

Here the quotation element (`<q>` . . . `</q>`) is nested inside the emphasis element (`<em>` . . . `</em>`). Moreover the "break" element (`<br />`), which usually causes subsequent text to start on a new line, is explicitly a singleton; it has a slash before its closing "`>`" delimiter. (To avoid confusing certain legacy browsers, one sometimes needs a space in front of the slash.) The example fragment would be malformed if the slash were missing, if the opening `<em><q>` tags were reversed (`<q><em>`), or if the attribute value `"favorite"` had not be enclosed in quote marks. An HTML parser would tolerate these errors; an XML parser will not. ■

The set of tags to be used in an XML document can be specified by naming a *document type definition* (DTD) in the document's `DOCTYPE` header, or by naming an *XML Schema* in an attribute of the document's top-level tag. (XML Schemas are a newer format, but DTDs remain in widespread use.) Among other things, a DTD or Schema indicates which tags are allowed, whether those tags are pairs or singletons, whether they permit attributes, and whether any attributes are mandatory. If a document has no DTD or Schema, it is said to define a DTD *implicitly* by

DESIGN & IMPLEMENTATION

14.14 W3C and WHATWG

Standardization efforts for HTML have a complicated history. With the completion in 1998 of the XML 1.0 specification, the World Wide Web Consortium (W3C) focused on XHTML, in an effort to push the world toward a "cleaned-up," XML-compliant version of HTML. Over the next few years, this strategy proved increasingly contentious. In 2004, a group of influential individuals from Apple, Mozilla, and Opera split off to form a separate Web Hypertext Application Technology Working Group (WHATWG), with the goal of evolving HTML in a way that preserved complete backward compatibility and interoperability. In 2006, the W3C reconsidered its position, and began to work with WHATWG toward what eventually became HTML5.

As of 2015, W3C and WHATWG remain separate organizations. Their standards, while very similar, are not entirely compatible. Both groups acknowledge that the world would be best served by a uniform definition of HTML, but their approaches to standardization differ greatly. W3C develops dated, numbered documents that codify the notion of conformance with a particular version of the standard. WHATWG maintains an unnumbered "living standard" that evolves continuously over time. W3C is more willing to label certain practice as noncomforming; WHATWG believes that its standard should reflect actual practice, as implemented in all past browsers by all major vendors. Both groups distinguish carefully between what a conforming document should contain and what a conforming browser should be able to render: the latter is significant superset of the former.

```
<!DOCTYPE html>
<html xmlns="http://www.w3.org/1999/xhtml" lang="en" xml:lang="en">
<head>
<meta charset="UTF-8" />
<title>Favorite Quote</title>
</head>
<body>
<p>
<em><q id="favorite">I defy
the tyranny of precedent</q></em><br />
(Clara Barton)
</p>
</body>
</html>
```

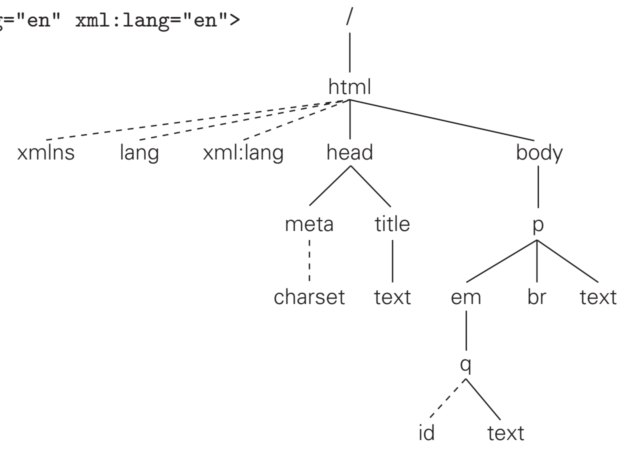


Figure 14.22 **A complete XHTML document and its corresponding tree.** Child elements are shown with solid lines, attributes with dashed lines.

virtue of which tags are actually used. Implicit definition suffices for the examples in this chapter.

EXAMPLE 14.83
XHTML to display a favorite quote

Because tags must nest in XML, a document has a natural tree-based structure. Figure C-14.22 shows the source for a small but complete polyglot HTML5 document, together with the tree it represents. There are three kinds of nodes in the tree: elements (delimited by tags in the source), text, and attributes. The internal (nonleaf) nodes are all elements. Everything nested between the beginning and ending tags of an element is an attribute or child of that element in the tree.

The root of our document, named "/" by convention, has one child—the `html` element. This in turn has three attributes—`xmlns`, `lang`, and `xml:lang`—and two child elements—`head` and `body`. The `xmlns` attribute specifies a URI for our document's *namespace*. This serves a purpose similar to that of C++ namespaces or Java packages (Section 3.8): it allows us to give tag names a disambiguating prefix: `xhtml:table` versus `furniture:table`. With the value we have specified for the `xmlns` attribute, any tag in the document that doesn't have a prefix will automatically be interpreted as being in the `xhtml` namespace. The `lang` and `xml:lang` tags specify the source language (English) for HTML and XML parsers, respectively. ■

XSLT and XPath

XSL (extensible stylesheet language) can be thought of as a language for specifying what to *do* with an XML document. It has four sublanguages, called XSLT, XPath, XSL-FO, and XQuery. XSLT is a scripting language that takes XML as input and produces textual output—often transformed XML or HTML, but potentially other formats as well.

EXAMPLE 14.84
XPath names for XHTML elements

XPath is a language used to name things in XML documents. XPath names frequently appear in the attributes of XSLT elements. Returning to Figure C-14.22, the quotation element of our document could be named in XPath as `/html/body/p/em/q`. The emphasis element and its break and text-node siblings, together, could be named as `/html/body/p/*`. XPath includes a rich set of naming mechanisms, including absolute (from the root) and relative (from the current node) navigation, wildcards, predicates, substring and regular expression manipulation, and counting and arithmetic functions. We will see some of these in the extended example below. ■

XSL-FO (XSL formatting objects) is a set of tags to specify the *layout* (presentation) of a document, in terms of pages, regions (e.g., header, body, footer), blocks (paragraph, table, list), lines, and in-line elements (character, image). An XSLT script might be used to add XSL-FO tags to an XML document, or to transform a document that already has XSL-FO tags in it—perhaps to split a long single-page document intended for the Web into a multipage document intended for printing on paper.

XQuery is a language in which to frame information-retrieval questions for a database stored in XML format. (In a bibliographic database, for example, we might use XQuery look for journal articles written since the turn of the century.) The purpose and behavior of XQuery parallel those of SQL, the standard language used for relational database queries. For the sake of simplicity, we will not use XSL-FO or XQuery in our extended example. Rather we will peruse an entire XML document, using XSLT to format its content as HTML.

An XML document can explicitly specify an XSLT script that should be used to transform or format it. All major browsers today include an XSLT interpreter, and will perform the transformation on the client machine. This is a standard but somewhat restrictive way to go about things: by tying a single stylesheet to the XML file we compromise the separation between content and presentation that was a principal motivation for creating XML in the first place. An alternative is to use client-side JavaScript or server-side PHP to invoke the XSLT processor, passing the XML document and the XSLT script as arguments.

Extended Example: Bibliographic Formatting

EXAMPLE 14.85
Creating a reference list with XSLT

As an example of a task for which we might realistically use XSLT, consider the creation of a bibliographic reference list. Figure C-14.23 contains XML source for such a list. (Field names have been borrowed from BIBTEX [Lam94, App. B].) The document begins with a declaration to specify the XML version and character encoding, and a processing instruction to specify the XSL stylesheet to be used to format the file. These declarations are included for the benefit of tools that process the document; they aren't part of the XML source itself. (Note the syntactic resemblance to the *processing instructions* used in Section 14.3.2 to provide input to the PHP interpreter.)

At the top level, the `bibliography` element consists of a series of `book`, `article`, and `inproceedings` elements, each of which may contain elements for author and editor names, title, publisher, date and address, and so on. Some

```
<?xml version="1.0" encoding="UTF-8"?>
<?xml-stylesheet type="text/xsl" href="bib.xsl"?>
<bibliography>
  <book>
    <author>Guido van Rossum</author>
    <editor>Fred L. Drake, Jr.</editor>
    <title>The Python Language Reference Manual (version 3.2)</title>
    <publisher>Network Theory, Ltd.</publisher>
    <address>Bristol, UK</address>
    <year>2011</year>
    <note>Available at <uri>http://www.network-theory.co.uk/python/language/</uri></note>
  </book>
  <article>
    <author>John K. Ousterhout</author>
    <title>Scripting: Higher-Level Programming for the 21st Century</title>
    <journal>Computer</journal>
    <volume>31</volume>
    <number>3</number>
    <month>March</month>
    <year>1998</year>
    <pages>23&#8211;30</pages>
  </article>
  <inproceedings>
    <author>Theodor Holm Nelson</author>
    <title>Complex Information Processing: A File Structure for the
        Complex, the Changing, and the Indeterminate</title>
    <booktitle>Proceedings of the Twentieth ACM National Conference</booktitle>
    <month>August</month>
    <year>1965</year>
    <address>Cleveland, OH</address>
    <pages>84&#8211;100</pages>
  </inproceedings>
  <inproceedings>
    <author>Stephan Kepser</author>
    <title>A Simple Proof for the Turing-Completeness of XSLT and XQuery</title>
    <booktitle>Proceedings, Extreme Markup Languages 2004</booktitle>
    <address>Montr&#233;al, Canada</address>
    <year>2004</year>
    <month>August</month>
    <note>Available at <uri>http://conferences.idealliance.org/extreme/html/2004/Kepser01/
        EML2004Kepser01.html</uri></note>
  </inproceedings>
```

Figure 14.23 **A bibliography in XML.** References (two books, a journal article, and three conference papers) appear in arbitrary order. The Kepser URI has been wrapped to fit on the printed page. *(continued)*

```
  <inproceedings>
    <author>David G. Korn</author>
    <title><code>ksh</code>: An Extensible High Level Language</title>
    <booktitle>Proceedings of the USENIX Very High Level Languages Symposium</booktitle>
    <address>Santa Fe, NM</address>
    <year>1994</year>
    <month>October</month>
    <pages>129&#8211;146</pages>
  </inproceedings>
  <book>
    <author>Tom Christiansen</author>
    <author>brian d foy</author>
    <author>Larry Wall</author>
    <author>Jon Orwant</author>
    <title>Programming Perl</title>
    <edition>fourth</edition>
    <publisher>O&#8217;Reilly Media</publisher>
    <address>Sebastopol, CA</address>
    <year>2012</year>
  </book>
</bibliography>
```

Figure 14.23 *(continued)*

elements may contain nested `uri` elements, which specify on-line links. Characters that cannot be represented in ASCII are shown as Unicode *character entities*, as described in Sidebar 7.3.

Figure C-14.24 contains an XSLT stylesheet (script) to format the bibliography as HTML, which may then be rendered in a browser. This script was named at the beginning of the XML document (Figure C-14.23). In a manner analogous to that of the XML document, the script begins with a declaration to specify the XML version and character encoding, and an `xsl:stylesheet` element to specify the XSL version and namespace. The remainder of the script contains a mix of XSL and HTML elements. The XSL tags all specify the `xsl:` namespace explicitly. They are recognized by the XSLT processor. Elements from other namespaces are treated as ordinary text, to be copied through to the output when encountered.

The fundamental construct in XSLT is the `template`, which specifies a set of *instructions* to be applied to nodes in an XML source tree. Templates are typically invoked by executing an `apply-templates` or a `call-template` instruction in some other template. Each invocation has a concept of *current node*. The execution as a whole begins by invoking an initial template with the root of the source tree (`/`) as current node. In our bibliographic example, the initial template is the one at the top of the script, because its `match` attribute is the XPath expression `"/"`. The body of the initial template begins with a string of HTML elements and text. This string is copied directly to the output. The `for-each` element, however, is an XSLT instruction, so it is executed.

```
<?xml version="1.0" encoding="UTF-8"?>
<xsl:stylesheet version="1.0" xmlns:xsl="http://www.w3.org/1999/XSL/Transform">

<xsl:template match="/">
  <html><head><title>Bibliography</title></head><body><h1>Bibliography</h1><ol>
    <xsl:for-each select="bibliography/*"><xsl:sort select="title"/>
      <li><xsl:apply-templates select="."/></li>
    </xsl:for-each>
  </ol></body></html>
</xsl:template>

<xsl:template match="bibliography/article">
  <q><xsl:apply-templates select="title/node()"/>,</q>
  by <xsl:call-template name="author-list"/>.&#160;
  <em><xsl:apply-templates select="journal/node()"/>
  <xsl:text> </xsl:text><xsl:apply-templates select="volume/node()"/>
  </em>:<xsl:apply-templates select="number/node()"/>
  (<xsl:apply-templates select="month/node()"/><xsl:text> </xsl:text>
    <xsl:apply-templates select="year/node()"/>),
  pages <xsl:apply-templates select="pages/node()"/>.
  <xsl:if test="note"><xsl:apply-templates select="note/node()"/>.</xsl:if>
</xsl:template>

<xsl:template match="bibliography/book">
  <em><xsl:apply-templates select="title/node()"/>,</em>
  by <xsl:call-template name="author-list"/>.&#160;
  <xsl:apply-templates select="publisher/node()"/>,
  <xsl:apply-templates select="address/node()"/>,
  <xsl:if test="edition">
    <xsl:apply-templates select="edition/node()"/> edition, </xsl:if>
  <xsl:apply-templates select="year/node()"/>.
  <xsl:if test="note"><xsl:apply-templates select="note/node()"/>.</xsl:if>
</xsl:template>

<xsl:template match="bibliography/inproceedings">
  <q><xsl:apply-templates select="title/node()"/>,</q>
  by <xsl:call-template name="author-list"/>.&#160;
  In <em><xsl:apply-templates select="booktitle/node()"/></em>
  <xsl:if test="pages">, pages <xsl:apply-templates select="pages/node()"/></xsl:if>
  <xsl:if test="address">, <xsl:apply-templates select="address/node()"/></xsl:if>
  <xsl:if test="month">, <xsl:apply-templates select="month/node()"/></xsl:if>
  <xsl:if test="year">, <xsl:apply-templates select="year/node()"/></xsl:if>.
  <xsl:if test="note"><xsl:apply-templates select="note/node()"/>.</xsl:if>
</xsl:template>
```

Figure 14.24 Bibliography stylesheet in XSL. This script will generate HTML when applied to a bibliography like that of Figure C-14.23. *(continued)*

```
<xsl:template name="author-list">       <!-- format author list -->
  <xsl:for-each select="author|editor">
    <xsl:if test="last() > 1 and position() = last()"> and </xsl:if>
    <xsl:apply-templates select="./node()"/>
    <xsl:if test="self::editor"> (editor)</xsl:if>
    <xsl:if test="last() > 2 and last() > position()">, </xsl:if>
  </xsl:for-each>
</xsl:template>

<xsl:template match="uri">              <!-- format link -->
    <a><xsl:attribute name="href"><xsl:value-of select="."/></xsl:attribute>
    <xsl:value-of select="substring-after(., 'http://')"/></a>
</xsl:template>

<xsl:template match="@*|node()">        <!-- default: copy content -->
  <xsl:copy><xsl:apply-templates select="@*|node()"/></xsl:copy>
</xsl:template>

</xsl:stylesheet>
```

Figure 14.24 *(continued)*

The `select` attribute of the `for-each` element uses an XPath expression (`"bibliography/*"`) to build a *node set* consisting of all top-level entries in our bibliography. Other expressions could have been used if we wanted to be selective: `"bibliography/*[year>=2000]"` would match only recent entries; `"bibliography/*[note]"` would match only entries with `note` elements; `"bibliography/article|bibliography/book"` would match only articles and books.

The nested `sort` instruction forces the selected node set to be ordered alphabetically by title. The body of the `for-each` is then executed with each entry in turn selected as current node. The body contains a recursive invocation of `apply-templates`, bracketed by HTML list tags (`<li>` ... `</li>`). These tags are copied to the output, with the result of the recursive call nested in between.

So how does the recursive call work? Its `select` attribute, like that of `for-each`, uses XPath to build a node set. In this case it is the trivial node set containing only `"."`, the current node of the current iteration of `for-each`. The XSLT processor searches for a template that matches this node. We have created three appropriate candidates, one for each kind of bibliographic entry. When it finds the matching template, the processor invokes it, with an updated notion of current node.

Each of our three main templates contains a set of instructions to format its kind of entry (article, book, conference paper). Most of the instructions use additional invocations of `apply-templates` to format individual portions of an entry (author, title, publisher, etc.). Interspersed in these instructions are snippets of text and HTML elements. In several cases we use an `if` instruction to generate output only when a given XML element is present in the source. In most of these the recursive call uses the XPath `node()` function to select all children of the element in question.

White space is ignored when it comes between the end of one instruction and the beginning of the next. To force white space into the output in this case, we must delimit it with `<text>`...`</text>` tags. Extra white space (e.g., after the ends of sentences) is specified with the "nonbreaking space" character entity, `&#160;`.

Three extra templates end our script. The most interesting of these serves to format author lists. It has a `name` attribute rather than a `match` attribute, and is invoked with `call-template` rather than `apply-templates`. A `called` template always takes the current node of the caller—in this case the node that represents a bibliographic entry. Internally, the author list template executes a `for-each` instruction that selects all child nodes representing authors or editors. The `for-each`, in turn, uses the XPath `last()` and `position()` functions to determine how many names there are, and where each name falls in the list. It inserts the word "and" between the final two names, and puts commas after all names but the last in lists of three or more.

The template with `match="uri"` serves to format URIs that appear anywhere in the XML source. It creates an HTML link in the output, but uses the XPath `substring-after` function to strip the leading *http://* off the visible text. XPath provides a variety of similar functions for string and regular expression manipulation. The `value-of` instruction copies the contents of the selected node to the output, as a character string.

Our final template serves as a default case. The XPath expression `"@*|node()"` will match any attribute or other node in the XML source. Inside, the `copy` instruction copies the node's tags, if any, to the output, with the result of a recursive call to `apply-templates` in between. The `"@*|node()"` on the recursive call selects a node set consisting of all the current node's attributes and children. The end result is that any XML elements in the source that are delimited by tags for which we do not have special templates will be regenerated in the output just as they appear in the source. The recursion stops at text nodes and attributes, which are the leaves of the XML tree.

HTML output from our script appears in Figure C-14.25. The rendered web page appears in Figure C-14.26.

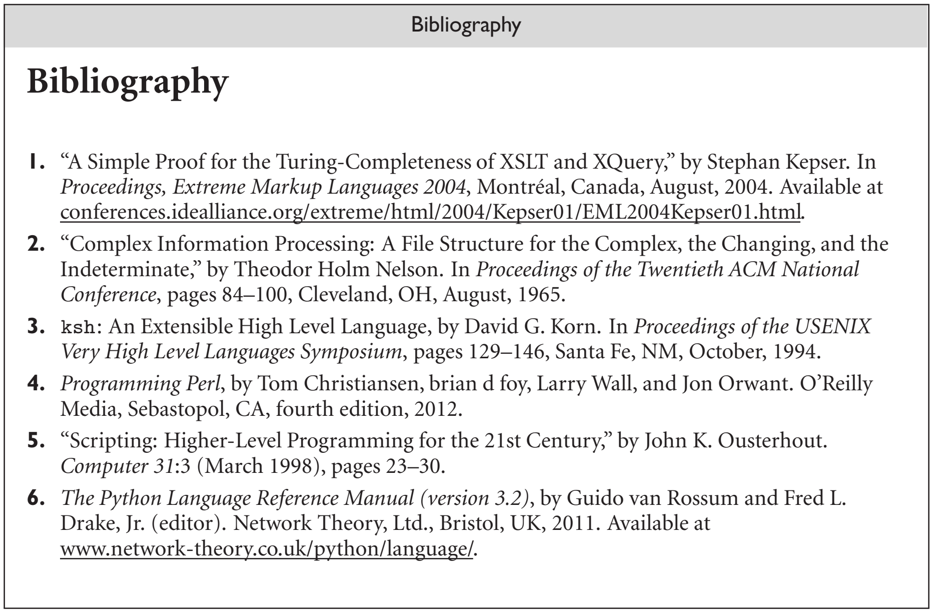
While lengthy by the standards of this text, our example illustrates only a fraction of the capabilities of XSLT. In the standard categorization of programming languages, the notation is strongly declarative: values may have names, but there are no mutable variables, and no side effects. There is a limited looping mechanism (`for-each`), but the real power comes from recursion, and from recursive traversal of XML trees in particular. ■

```
<html><head><title>Bibliography</title></head>
<body><h1>Bibliography</h1><ol>
<li>
  <q>A Simple Proof for the Turing-Completeness of XSLT and XQuery,</q>
  by Stephan Kepser.&nbsp; In <em>Proceedings, Extreme Markup Languages
  2004</em>, Montr&eacute;al, Canada, August, 2004.  Available at
  <a href="http://conferences.idealliance.org/extreme/html/2004/Kepser01
/EML2004Kepser01.html">conferences.idealliance.org/extreme/html/2004
/Kepser01/EML2004Kepser01.html</a>.</li>
<li>
  <q>Complex Information Processing: A File Structure for the Complex,
  the Changing, and the Indeterminate,</q> by Theodor Holm Nelson.&nbsp;
  In <em>Proceedings of the Twentieth ACM National Conference</em>,
  pages 84&ndash;100, Cleveland, OH, August, 1965.</li>
<li>
  <q><code>ksh</code>: An Extensible High Level Language,</q> by David
  G. Korn.&nbsp; In <em>Proceedings of the USENIX Very High Level Languages
  Symposium</em>, pages 129&ndash;146, Santa Fe, NM, October, 1994.</li>
<li>
  <em>Programming Perl,</em> by Tom Christiansen, brian d foy, Larry Wall,
  and Jon Orwant.&nbsp; O&rsquo;Reilly Media, Sebastopol, CA, fourth
  edition, 2012.</li>
<li>
  <q>Scripting: Higher-Level Programming for the 21st Century,</q> by
  John K. Ousterhout.&nbsp; <em>Computer 31</em>:3 (March 1998), pages
  23&ndash;30.</li>
<li>
  <em>The Python Language Reference Manual (version 3.2),</em> by Guido
  van Rossum and Fred L. Drake, Jr. (editor).&nbsp; Network Theory, Ltd.,
  Bristol, UK, 2011.  Available at <a href="http://www.network-theory.co.uk
/python/language/">www.network-theory.co.uk/python/language/</a>.</li>
</ol>
</body></html>
```

Figure 14.25 Result of applying the stylesheet of Figure C-14.24 to the bibliography of Figure C-14.23.

✓ CHECK YOUR UNDERSTANDING

55. Explain the relationships among SGML, HTML, and XML. What are their corresponding stylesheet languages?
56. Why does XML work so hard to distinguish between *content* and *presentation*?
57. What are the four main components of XSL? What are their respective purposes?
58. What is XHTML? How does it differ from "ordinary" HTML?
59. Explain the correspondence between XML documents and trees.
60. What does it mean for an XML document to be *well formed*?


Bibliography

Bibliography

1. “A Simple Proof for the Turing-Completeness of XSLT and XQuery,” by Stephan Kepser. In *Proceedings, Extreme Markup Languages 2004*, Montréal, Canada, August, 2004. Available at conferences.idealliance.org/extreme/html/2004/Kepser01/EML2004Kepser01.html.
2. “Complex Information Processing: A File Structure for the Complex, the Changing, and the Indeterminate,” by Theodor Holm Nelson. In *Proceedings of the Twentieth ACM National Conference*, pages 84–100, Cleveland, OH, August, 1965.
3. `ksh`: An Extensible High Level Language, by David G. Korn. In *Proceedings of the USENIX Very High Level Languages Symposium*, pages 129–146, Santa Fe, NM, October, 1994.
4. *Programming Perl*, by Tom Christiansen, brian d foy, Larry Wall, and Jon Orwant. O’Reilly Media, Sebastopol, CA, fourth edition, 2012.
5. “Scripting: Higher-Level Programming for the 21st Century,” by John K. Ousterhout. *Computer 31*:3 (March 1998), pages 23–30.
6. *The Python Language Reference Manual (version 3.2)*, by Guido van Rossum and Fred L. Drake, Jr. (editor). Network Theory, Ltd., Bristol, UK, 2011. Available at www.network-theory.co.uk/python/language/.

Figure 14.26 Rendered version of the HTML in Figure C-14.25.

61. Explain the distinctions (syntactic and semantic) among *elements*, *declarations*, and *processing instructions* in XML. Also explain the distinctions among *elements*, *tags*, and *attributes*.

62. Summarize the execution model of XSLT. In a nutshell, how does it work?

63. Explain the difference between *applying* templates and *calling* them in XSLT.

14 Scripting Languages

14.6 Exercises

14.19 Modify the XSLT of Figure C-14.24 to do one or more of the following:

(a) Alter the titles of conference papers so that only first words, words that follow a dash or colon (and thus begin a subtitle), and proper nouns are capitalized. You will need to adopt a convention by which the creator of the document can identify proper nouns.

(b) Sort entries by the last name of the first author or editor. You will need to adopt a convention by which the creator of the document can identify compound last names ("von Neumann," for example, should be alphabetized under 'v').

(c) Allow bibliographic entries to contain an `abstract` element, which when formatted appears as an indented block of text in a smaller font.

(d) In addition to the `book`, `article`, and `inproceedings` elements, add support for other kinds of entries, such as manuals, technical reports, theses, newspaper articles, web sites, and so on. You may want to draw inspiration from the categories supported by BIBTEX [Lam94, App. B].

(e) Format entries according to some standard style convention (e.g., that of the Chicago Manual of Style [*www.chicagomanualofstyle.org/16/ch14/ch14_toc.html*] or the ACM Transactions [*www.acm.org/publications/article-templates/acm-latex-style-guide*]).

14.20 Suppose bibliographic entries in Figure C-14.23 contain a mandatory `key` element, and that other documents can contain matching `cite` elements. Create an XSLT script that imitates the work of BibTEX. Your script should

(a) read an XML document, find all the `cite` elements, collect the keys they contain, and replace them with `bibref` elements that contain small integers instead.

(b) read a separate XML bibliography document, extract the entries with matching keys, and write them, in sorted order, to a new (and probably smaller) bibliography.

The small numbers in the `bibref` elements of the new document from (a) should match the corresponding numbered entries in the new bibliography from (b).

14.21 Write a program that will read an XHTML file and print an outline of its contents, by extracting all `<title>`, `<h1>`, `<h2>`, and `<h3>` elements, and printing them at varying levels of indentation. Write

(a) in C or Java

(b) in `sed` or `awk`

(c) in Perl, Python, Tcl, or Ruby

(d) in XSLT

Compare and contrast your solutions.

14 Scripting Languages

14.7 Explorations

14.32 Learn more about DTDs and XML Schemas. Compare the DTD and XML Schema definitions of XHTML. What appear to the prospects for migrating to the newer specification language?

14.33 Academics often keep lists of publications in multiple places and formats: an on-line web page, a printable resume, a BIBTEX database for paper writing [Lam94, App. B]. Using XSLT, build a set of tools that will construct these lists automatically from a single XML source file.

14.34 Learn about XSL-FO. Use it to reimplement Example C-14.85. Your new version should be a two-stage process: one XSLT script should add formatting tags to the XML bibliography; a second should convert the tagged bibliography to XHTML. Try to make these stages as general as possible: you should be able to modify the appearance of the output list by changing the first script only. You should also be able to write alternative versions of the second script that generate output in formats other than XHTML (e.g., LaTeX).

14.35 Learn more about the history of W3C and WHATWG. What are the comparative advantages and disadvantages of their approaches to standardization? Do you find yourself more in sympathy with one approach or the other? How large are the technical differences between the most recent versions of the HTML standards? Are these differences significant enough to pose a problem for web developers?

15

Building a Runnable Program

15.2.1 GIMPLE and RTL

Traditionally, all machine-independent code improvement in `gcc` was based on RTL. Over time it became clear that the IF had become an obstacle to further improvements in the compiler, and that a higher-level form was needed. GIMPLE was introduced to meet that need. As of `gcc` v.4.9, GENERIC is used for semantic analysis and, in a few cases, for certain language-specific code improvement. As its final task, each front end converts the program from GENERIC into GIMPLE. Depending on the requested level of code improvement, the "middle end" may perform over 140 phases of code improvement and transformation on the GIMPLE representation, after which it converts to RTL and performs as many as 70 additional phases before handing the result to the back end for target code generation.

Both GIMPLE and RTL are meant to be kept in memory across compiler phases, rather than being written to a file. Both IFs have a human-readable external format, which the compiler can write and (partially) read, but this format is not needed by the compiler: the internal version is much better suited for automatic manipulation.

GIMPLE

The GIMPLE code generated by a `gcc` front end is essentially a distillation of GENERIC, with many of the most complex (and often language-specific) features "lowered" into a smaller, common set of tree node types. As a simple example, consider the `gcd` program of Example 1.20:

EXAMPLE 15.19
GCD program in GIMPLE

```
int main () {
    int i = getint();
    int j = getint();
    while (i != j) {
        if (i > j) i = i - j;
        else j = j - i;
    }
    putint(i);
}
```

Figure C-15.11 illustrates the "high GIMPLE" produced by the C front end of gcc 4.8 when given this program as input. If we compare this GIMPLE code to Figure 15.2, which loosely[1] resembles GENERIC, we see at least two significant differences. First, all of the nodes that comprise a subroutine appear on a single list, with control flow represented by explicit gotos and by true and false branches for conditions. Second, both conditions and assignments have been designed to capture an embedded binary expression, allowing us in many cases to collapse a small subtree into a GIMPLE single node. ■

Over the course of its many phases, the gcc middle end will make many additional changes to this code, not only to improve its quality but also to further lower its level of abstraction. This "flattening" of the tree makes it easier to translate into RTL.

Perhaps the most significant transformation of GIMPLE is the conversion to *static single assignment (SSA) form*. We will study SSA in more detail in Section C-17.4.1. Briefly, the SSA conversion introduces extra variable names into the program in such a way that nothing is ever written in more than one place. If there are 10 assignments to variable foo in the source code, there will be (at least) 10 separate variables $\text{foo}_1, \ldots, \text{foo}_{10}$ in SSA. When control paths merge (e.g., after an if... then ... else), versions of a variable arriving on different paths are combined, using a hypothetical "phi function" to create yet another version ($\text{foo}_{11} := \phi(\text{foo}_1, \text{foo}_2)$). As in functional programming languages, the single-assignment character of SSA means that expressions are *referentially transparent*—independent of evaluation order. Referential transparency significantly simplifies many forms of code improvement.

RTL

RTL is loosely based on the S-expressions of Lisp. Each RTL expression consists of an operator or expression type and a sequence of operands. In its external form, these are represented by a parenthesized list in which the element immediately inside the left parenthesis is the operator. Each such list is then embedded

1 Unlike the informal notation of Figure 15.2, GENERIC and GIMPLE make no distinction between syntax tree nodes and symbol table nodes. In effect, the symbol table is merged into the syntax tree.

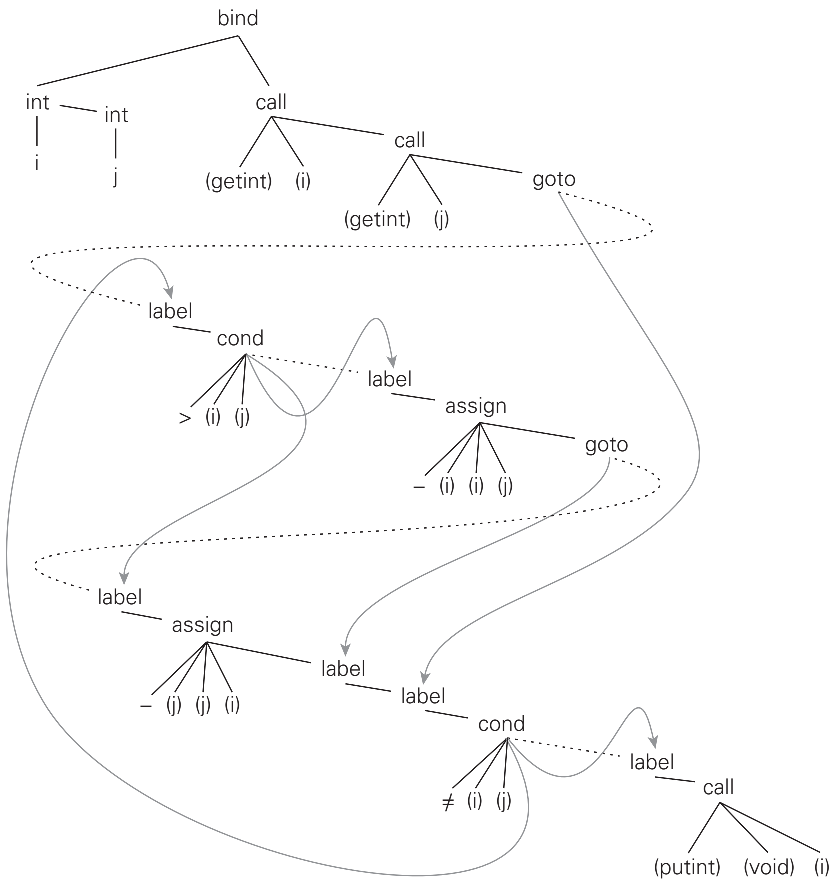


Figure 15.11 **Simplified GIMPLE for the `gcd` program.** The left child of the `bind` node holds local symbol table information; references to this information—and to global functions getint and putint—are indicated in the rest of the figure with parenthesized names. The first child of a call node names the function, the second the place to assign the return value, and the rest the arguments. An assign node with children ⟨op, a, b, c⟩ represents the assignment a := b op c. In each condition node, the first three children are a comparison operator and its operands; the last two are pointers to the subtrees for the outcomes true and false.

in a wrapper that points to predecessor and successor expressions in linear order. Internally, RTL expressions are represented by C structs and pointers. This pointer-rich structure constitutes the interface among the compiler's many back-end phases. There are several dozen expression types, including constants, references to values in memory or registers, arithmetic and logical operations, comparisons, bit-field manipulations, type conversions, and stores to memory or registers.

The body of a subroutine consists of a sequence of RTL expressions. Each expression in the sequence is called an `insn` (instruction). Each `insn` begins with one of six special codes:

insn: an "ordinary" RTL expression.

jump_insn: an expression that may transfer control to a label.

call_insn: an expression that may make a subroutine call.

code_label: a possible target of a jump.

barrier: an indication that the previous `insn` always jumps away. Control will never "fall through" to here.

note: a pure annotation. There are nine different kinds of these, to identify the tops and bottoms of loops, scopes, subroutines, and so on.

The sequence is not always completely linear; `insns` are sometimes collected into pairs or triples that correspond to target machine instructions with delay slots. Over a dozen different kinds of (non-*note*) annotations can be attached to an individual `insn`, to identify side effects, specify target machine instructions or registers, keep track of the points at which values are defined and used, automatically increment or decrement registers that are used to iterate over an array, and so on. `Insns` may also refer to various dynamically allocated structures, including the symbol table.

EXAMPLE 15.20
An RTL `insn` sequence

A simplified `insn` sequence for the code of Example C-15.19 appears in Figure C-15.12. The three leading numbers in each `insn` represent the `insn`'s unique id and those of its predecessor and successor, respectively. The fourth, when present, identifies the `insn`'s basic block. Fields for the various `insn` annotations are not shown. The `:SI` *mode specifier* on a memory or register reference indicates access to a single (4-byte) integer; `:DI` and `:QI` modes correspond to double (8-byte) and quarter (1-byte) integers.

A full explanation of the RTL notation is beyond what we can cover here. As an example, however, `insn` 26 loads the memory location found 4 bytes back from the frame pointer (namely, `i`) into virtual register 64. The following `insn`, 27, sets the memory location found 8 bytes back from the frame pointer (namely, `j`) to the result of subtracting register 64 from that same memory location. In parallel (as a side effect), `insn` 27 also "clobbers" (overwrites) the contents of virtual condition code register 17. ■

In order to generate target code, the back end matches `insns` against patterns stored in a semiformal description of the target machine. Both this description and the routines that manipulate the machine-dependent parts of an `insn` are segregated into a relatively small number of separately compiled files. As a result, much of the compiler back end is machine independent, and need not actually be modified when porting to a new machine.

```
(insn 5 2 6 2 (set (reg:QI 0 ax) (const_int 0)))
(call_insn 6 5 7 2 (set (reg:SI 0 ax) (call (mem:QI (symbol_ref:DI ("getint"))) (const_int 0))))
(insn 7 6 8 2 (set (reg:SI 60) (reg:SI 0 ax)))
(insn 8 7 9 2 (set (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4))) (reg:SI 60)))
(insn 9 8 10 2 (set (reg:QI 0 ax) (const_int 0)))
(call_insn 10 9 11 2 (set (reg:SI 0 ax) (call (mem:QI (symbol_ref:DI ("getint"))) (const_int 0))))
(insn 11 10 12 2 (set (reg:SI 61) (reg:SI 0 ax)))
(insn 12 11 13 2 (set (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -8))) (reg:SI 61)))
(jump_insn 13 12 14 2 (set (pc) (label_ref 28)))
(barrier 14 13 30)
(code_label 30 14 15 4 4 "")
(insn 16 15 17 4 (set (reg:SI 62) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4)))))
(insn 17 16 18 4 (set (reg:CCGC 17)
                 (compare:CCGC (reg:SI 62) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -8))))))
(jump_insn 18 17 19 4 (set (pc) (if_then_else (le (reg:CCGC 17) (const_int 0)) (label_ref 24) (pc))))
(insn 20 19 21 5 (set (reg:SI 63) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -8)))))
(insn 21 20 22 5 (parallel [
    (set (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4)))
         (minus:SI (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4))) (reg:SI 63)))
    (clobber (reg:CC 17))
]))
(jump_insn 22 21 23 5 (set (pc) (label_ref 28)))
(barrier 23 22 24)
(code_label 24 23 25 6 3 "")
(insn 26 25 27 6 (set (reg:SI 64) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4)))))
(insn 27 26 28 6 (parallel [
    (set (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -8)))
         (minus:SI (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -8))) (reg:SI 64)))
    (clobber (reg:CC 17))
]))
(code_label 28 27 29 7 2 "")
(insn 31 29 32 7 (set (reg:SI 65) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4)))))
(insn 32 31 33 7 (set (reg:CCZ 17)
                      (compare:CCZ (reg:SI 65) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -8))))))
(jump_insn 33 32 34 7 (set (pc) (if_then_else (ne (reg:CCZ 17) (const_int 0)) (label_ref 30) (pc))))
(insn 35 34 36 8 (set (reg:SI 66) (mem/c:SI (plus:DI (reg/f:DI 54) (const_int -4)))))
(insn 36 35 37 8 (set (reg:SI 5 di) (reg:SI 66)))
(call_insn 37 36 40 8 (call (mem:QI (symbol_ref:DI ("putint"))) (const_int 0)))
(insn 40 37 41 8 (clobber (reg/i:SI 0 ax)))
(insn 41 40 39 8 (clobber (reg:SI 59 [ <retval> ])))
(insn 39 41 42 8 (set (reg/i:SI 0 ax) (reg:SI 59 [ <retval> ])))
(insn 42 39 0 8 (use (reg/i:SI 0 ax)))
```

Figure 15.12 **Simplified textual RTL for the `gcd` program.** Most annotations (more than half the original length) have been elided here. Register 54 is the frame pointer. Local variable `i` is at offset -4. Local variable `j` is at offset -8.

✓ CHECK YOUR UNDERSTANDING

25. Characterize GIMPLE, RTL, Java bytecode, and Common Intermediate Language as high-, medium-, or low-level intermediate forms.
26. Name three languages (other than C) for which there exist `gcc` front ends.
27. What is the internal IF of `gcc`'s front ends?
28. Give brief descriptions of GIMPLE and RTL. How do they differ? Why was GIMPLE introduced?

15 Building a Runnable Program

15.7 Dynamic Linking

To be amenable to dynamic linking, a library must either (1) be located at the same address in every program that uses it, or (2) have no relocatable words in its code segment, so that the content of the segment does not depend on its address. The first approach is straightforward but restrictive: it generally requires that we assign a unique address to every sharable library; otherwise we run the risk that some newly created program will want to use two libraries that have been given overlapping address ranges. In Unix System V R3, which took the unique-address approach, shared libraries could only be installed by the system administrator. This requirement tended to limit the use of dynamic linking to a relatively small number of popular libraries. The second approach, in which a shared library can be linked at any address, requires the generation of *position-independent code*. It allows users to employ dynamic linking whenever they want, without administrator intervention.

The cost of user-managed dynamic linking is that executable programs are no longer self-contained. They depend for correct execution on the availability of appropriate dynamic libraries at execution time. If different programs are built with different expectations of (which versions of) which libraries will be available, conflicts can arise. On Microsoft platforms, where dynamic libraries have names ending in `.dll`, compatibility problems are sometimes referred to as "DLL hell." The frequency and severity of the problem can be minimized with good software engineering practice. In particular, a *package management system* may maintain a database of dependences between programs and libraries, and among the libraries themselves. If installer programs use the database correctly, problems will be detected at install time, when they can reasonably be addressed, rather than at the arbitrarily delayed point at which a program first attempts to use an incompatible or missing library.

15.7.1 Position-Independent Code

A code segment that contains no relocatable words is said to constitute *position-independent code* (PIC). To generate PIC, the compiler must observe the following rules:

1. Use PC-relative addressing, rather than jumps to absolute addresses, for all internal branches.
2. Similarly, avoid absolute references to statically allocated data, by using displacement addressing with respect to some standard base register. If the code and data segments are guaranteed to lie at a known offset from one another, then the program counter can be used for this purpose. Otherwise, the caller must initialize some other base register as part of the entry point's calling sequence.
3. Use an extra level of indirection for every control transfer out of the PIC segment, and for every load or store of static memory outside the corresponding data segment. The indirection allows the (non-PIC) target address to be kept in the data segment, which is private to each program instance.

EXAMPLE 15.21
PIC under x86/Linux

Exact details vary among processors, vendors, and operating systems. Conventions for `gcc` on recent versions of x86 Linux are illustrated in Figure C-15.13. Each code segment is accompanied by a *linkage table*—known in Linux as the segment's *global offset table* (GOT). This table lists the locations of all code and data whose addresses were not statically determined. All processes that use the same library share a single copy of the library's code segment, but each process has its own copy of the library's GOT. Both the code segment and the GOT can lie at different locations in the address spaces of different processes, but the *offset* between the two must always be the same.

Like the main program, each shared library is typically composed of multiple compilation units, joined together by a *static linker*, which resolves internal references. Resolution of references from the main program into shared libraries—or among the libraries themselves—is delayed until load time or run time, and is the job of the *dynamic linker*. By construction, shared libraries never make references back into the main program.

Libraries are permitted to have (process-private) data as well as code, but the total amount of such data is assumed (in Linux, at least) to be small enough that the data can be statically linked without wasting significant space. Each process therefore has a single data segment (shown in the figure at the lower left), containing the data of the main program and of all the libraries it may call, directly or indirectly. (Extensions to delay the linking of library data are considered in Exercise C-15.14.)

Focusing for the moment on the dashed arrows of the figure (and ignoring the dotted arrows), a read of `X` or `Y` in `main` can use a statically resolved address. A read of `X` or `Y` in `foo` uses PC-relative addressing to find the appropriate slot in `foo`'s GOT, and then loads `X` or `Y` indirectly. Similar indirection is required

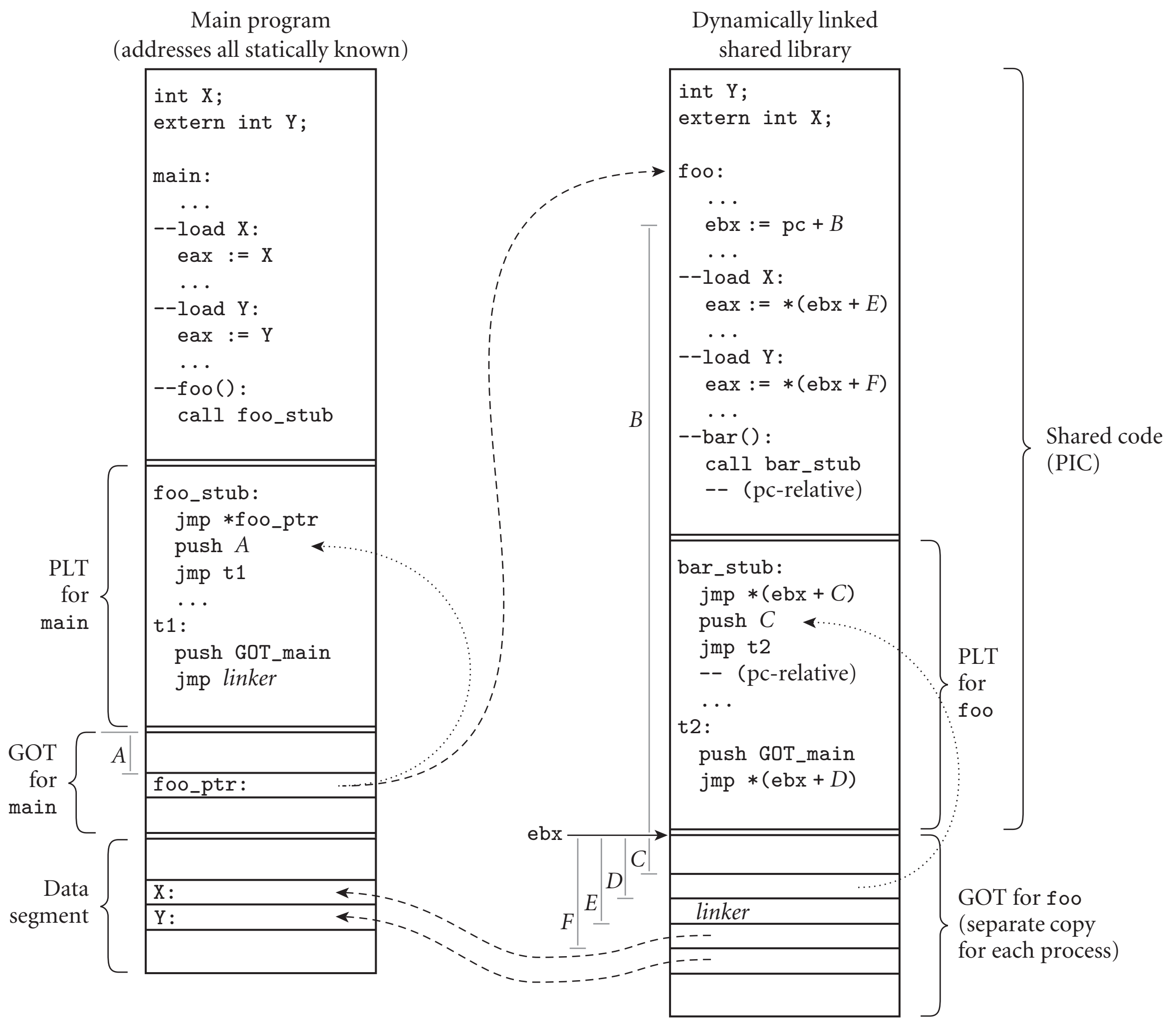


Figure 15.13 A dynamically linked shared library. Calls to foo and bar are made indirectly, using an address stored in the global offset tables (GOTs) of main and foo, respectively. Similarly, references to global variables X and Y, when made from foo, must employ a level of indirection. Resolved values are shown with dashed lines; initial values to support lazy linking (Section C-15.7.2) are shown with dotted lines. In the prologue of foo, register ebx is set to point to foo's GOT, using pc-relative arithmetic.

for subroutine calls into dynamically linked libraries. To avoid duplication of the indirection code, the compiler incorporates a (shared, read-only) `procedure linkage table` (PLT) in each code segment. To effect a call to `foo`, `main` calls a *stub* routine, here named `foo_stub`. This, in turn, performs an indirect jump to the address of `foo` found in `main`'s GOT. Inside `foo`, the call to `bar` is only slightly more complicated: the compiler must use PC-relative addressing to find the appropriate slot in foo's GOT. ■

EXAMPLE 15.22
PC-relative addressing on the x86

Most machines—including the x86—can perform branches and calls using PC-relative addressing. In our Linux example (Figure C-15.13), the machine-language encoding of `call bar_stub` in library `foo` will specify the offset between the call instruction and the `bar_stub` location in `foo`'s PLT.

Many machines can also use PC-relative addressing in load and store instructions. On the x86-64, for example, the load of `X` in `foo` could say `rax := *(rip + G)`, where *G* is the offset from the load instruction to `X`'s entry in `foo`'s GOT (on the x86-64, `rip` [instruction pointer register] is the name of the program counter). Unfortunately, the x86-32 does not support PC-relative addressing for loads and stores. To compensate, each PIC code segment on x86-32 Linux defines the following tiny subroutine:

```
get_pc:
    ebx := *esp     -- load location referred to by esp
    ret                 -- i.e., the return address -- into ebx
```

Given this definition, the pseudo-instruction `ebx := pc + B` in Figure C-15.13 can be implemented as

```
    call get_pc
    ebx += B
```

after which `ebx` can be used as the base for displacement addressing within `foo`'s GOT. ■

15.7.2 Fully Dynamic (Lazy) Linking

If all or most of the symbols exported by a shared library were always referenced by the parent program, it might make sense to link the library in its entirety at load time. When the program began running, its GOTs would then appear as suggested by the dashed arrows in Figure C-15.13. Certain systems indeed work in this fashion. In any given execution of a program, however, there may be references to libraries that are not actually used, because the input data never cause execution to follow the code path(s) on which the references appear. If these "potentially unnecessary" references are numerous, we may avoid a significant amount of work by linking the library *lazily* on demand. Moreover even in a program that uses all its symbols, incremental lazy linking may improve the system's interactive responsiveness by allowing programs to begin execution faster. A language system that allows the dynamic creation or discovery of program components (e.g., as in Common Lisp or Java) must also use lazy linking to delay the resolution of external references in dynamically compiled components.

EXAMPLE 15.23
Dynamic linking in Linux on the x86

When a Linux program first starts running, the data entries in its GOTs are indeed initialized as previously discussed; all addresses are known, because data locations are statically assigned. Code entries in the GOTs, however, point back into the corresponding PLTs, as suggested by the dotted arrows in Figure C-15.13.

Now consider what happens when `main` calls `foo_stub`. The `foo_ptr` entry in `main`'s GOT points to the second instruction of `foo_stub`—immediately after the indirect jump. That jump, in other words, ends up targeting the very next instruction, as if it had not happened at all. The next instruction, for its part, pushes onto the stack the offset of `foo`'s entry in `main`'s GOT. It then jumps (using PC-relative addressing) to a special entry in the PLT. This entry in turn pushes the address of `main`'s GOT and jumps to the dynamic linker, whose address is statically known. The dynamic linker consults symbol table information found in `main`'s executable file. Specifically, it looks up the GOT address and offset that were passed to it on the stack and discovers that they correspond to `foo`. It chooses a place for `foo` in the process's address space, creates a (process-specific) `foo` GOT at the appropriate offset (using symbol table information from `foo`'s own object file), initializes any data locations in that GOT to point to appropriate locations in the process's data segment, and initializes code locations in the GOT to point to the second instructions of the corresponding entries in `foo`'s PLT.

Now that `foo` has been given a location in the process's address space, the dynamic linker can modify `foo`'s entry in `main`'s PLT so that subsequent calls from `main` to `foo` will skip the linking step, and instead follow the single indirection suggested by the dashed arrow in Figure C-15.13. Last of all, the linker pops its arguments from the stack (leaving the return address pushed by `main` in its original call to `foo_stub`) and branches directly to `foo`. When `foo` completes, it will return to the correct address in `main`.

If and when `foo` calls `bar`, a similar series of events will take place. The principal difference is that both the body of `foo` and the stubs in its PLT must use PC-relative addressing to access entries in `foo`'s GOT. ■

✓ CHECK YOUR UNDERSTANDING

29. Explain the addressing challenge faced by dynamic linking systems.

30. What is *position-independent code*? What is it good for? What special precautions must a compiler follow in order to produce it?

31. Explain the need for PC-relative addressing in position-independent code. How is it accomplished on the x86-32?

32. What is the purpose of a *linkage table*?

33. What is *lazy* dynamic linking? What is its purpose? How does it work?

15 Building a Runnable Program

15.9 Exercises

15.12 Compare and contrast GIMPLE with the notation we have been using for syntax tree attribute grammars (Section 4.6).

15.13 PC-relative branches on many processors are limited in range—they can only target locations within 2^k bytes of the current PC, for some k less than the wordsize of the machine. Explain how to generate position-independent code that needs to branch farther than this.

15.14 We have noted that Linux creates a single data segment containing all the static data of libraries that might be called (directly or indirectly) by a given program. The space required for this segment is usually not a problem: most libraries have little static data—often none at all. Suppose this were not the case. If we wanted to perform dynamic linking for modules with large amounts of per-module static data, how could we extend Linux's dynamic linking mechanisms to perform fully dynamic (lazy) linking not only of code, but also of data?

15.15 In Example C-9.61 we described how the GNU Ada Translator (`gnat`) for the x86 uses dynamically generated code to represent a subroutine closure. Explain how a similar technique could be used to simplify the mechanism of Figure C-15.13, if we were willing to modify code segments at run time.

15 Building a Runnable Program

15.10 Explorations

15.21 Find the on-line documentation for `gcc`, which explains both GIMPLE and RTL, and enumerates command-line flags that will cause the compiler to dump its intermediate forms to standard output. (Version 4.8.4 of the compiler supports 26 such flags for GIMPLE and 67 for RTL.) Using appropriate flags and a small but nontrivial input program, arrange for the compiler to dump several versions of both GIMPLE and RTL. Study the output and describe how it has been changed by the intervening code improvement phases.

15.22 Find out how linking works under your favorite non-Linux system. Can code be dynamically linked? Can (nonprivileged) users create shared libraries? How does the loader or dynamic linker determine which libraries a program will need? How does it locate their object code? If your compiler can generate both position-independent and non-position-independent code, how do the two compare in size and run-time efficiency?

15.23 Learn about *pointer swizzling* [Wil92a], originally developed to run programs on machines with insufficient virtual address space. Explain its connection to dynamic linking.

15.24 Learn about ASIS, the Ada Semantic Interface Specification. How does it improve on tools based on the earlier Diana notation? How does it work in `gnat`?

15.25 We have had occasion in several previous sections to refer to the LLVM compiler suite. Much of the early work on LLVM revolved around its low-level IF, from which the system takes its name (Low Level Virtual Machine). Learn about this IF. How does it compare to RTL?

16 Run-Time Program Management

16.1.2 The Common Language Infrastructure

Work on the system that became the Common Language Infrastructure (CLI) began at Microsoft Corporation in the late 1990s, and was able to benefit from experience with Java and the JVM, which were already well established. The most significant differences between the virtual machines, however, stem from Microsoft's emphasis on cross-language interoperability—an emphasis that predates the JVM by many years.

Growing out of earlier work on the DDE, OLE, COM, ActiveX, and DCOM projects, the beta version of .NET was released in 2000. In addition to a virtual machine, it includes libraries, servers, and tools for a wide variety of local and distributed services, including user interface management, database access, networking, and security. A specification for the virtual machine—the CLI—was standardized by ECMA in 2001 and by the ISO in 2003. The standard has been updated several times over the years; version 6 was released in June 2012 [Int12a].

Perhaps the most significant contribution of the CLI is the definition of a Common Type System (CTS) for all supported languages. Encompassing nearly everything described in Chapters 8 and 10 of this book, the CTS provides a superset of what any particular language needs, while requiring common semantics and implementation wherever the type systems of more than one language intersect. In addition to the CTS, the CLI defines a virtual machine architecture, the VES (Virtual Execution System); an instruction set for that machine, the CIL (Common Intermediate Language); and a portable file format for code and metadata, PE (Portable Executable) assemblies.

C# is in some sense the premier language for .NET, and was developed concurrently with it. Several dozen languages have been ported to the CLI, however, and several of these, including Visual Basic, C++, and JScript, are now in widespread use. Several interesting challenges for the CTS were raised by the development of F#, an ML descendant designed by Microsoft and introduced in 2005.

Thanks to the ECMA/ISO standard, it is possible for organizations other than Microsoft to build implementations of the CLI. The leading such implementation is the open-source Mono project, led by Xamarin, Inc. Mono runs on a wide variety of platforms, but tends to lag slightly behind .NET in the addition of new features. Outside Microsoft, Java and the JVM still dominate. Within Microsoft, most new development today employs C#. Microsoft calls its CLI implementation the Common Language Runtime (CLR); it refers to CIL as Microsoft Intermediate Language (MSIL).

Architecture and Comparison to the JVM

In many ways, the CLI resembles the JVM. Both systems define a multithreaded, stack-based virtual machine, with built-in support for garbage collection, exceptions, virtual method dispatch, and mix-in inheritance. Both represent programs using a platform-independent, self-descriptive, bytecode notation. For languages like C#, the CLI provides all the safety of the JVM, including definite assignment, strong typing, and protection against overflow or underflow of the operand stack.

The biggest contrasts between the JVM and CLI stem from the latter's support for multiple programming languages (the following is not a comprehensive list).

Richer Type System The Common Type System (discussed below) supports both value and reference variables of structured types (the JVM is limited to references). The CTS also has true multidimensional arrays (allocated, contiguously, as a single operation); function pointers; explicit support for generics; and the ability to enforce structural type equivalence.

DESIGN & IMPLEMENTATION

16.7 Assuming a just-in-time compiler

Like the JVM, the CLI has behavior defined in terms of an abstract virtual machine. Where Java's virtual machine may in practice be either interpreted or just-in-time compiled, however, the CLI was designed from the outset for just-in-time compilation. Several minor differences between the virtual machines reflect this difference in expected implementations. Arithmetic instructions in Java bytecode generally include an explicit indication of operand type: there are, for example, four separate opcodes for 32- and 64-bit integer and floating-point addition. In the CLI's Common Intermediate Language (CIL), there is only one `add` instruction: it figures out what to do based on the types of its operands. In type-safe code, of course, the type of every operand is statically known, and either a compiler or an interpreter can inspect the types of arguments and figure out what to do. The compiler, however, only has to do this once, at compile time; the interpreter has to do it every time it encounters the instruction. In a similar vein, slots in the local variable array of the CLI VES can be of arbitrary size, and are required to hold a value of a single, statically known type throughout the execution of the method. For the sake of space efficiency and rapid indexing, the JVM reserves exactly 32 bits for every slot (`long`s and `double`s take two consecutive slots), and a given slot can be used for values of different types at different points in time.

Richer Calling Mechanisms To facilitate the implementation of functional languages, the CLI provides explicit tail-recursive function calls (Section 6.6.1); these discard the caller's frame while retaining the dynamic link. The CLI also supports both value and reference parameters, variable numbers of parameters (in the fully general sense of C), multiple return values, and nonvirtual methods, all of which the JVM lacks.

Unsafe Code For the benefit of C, C++, and other non-type-safe languages, the CLI supports explicitly unsafe operations: nonconverting type casts, dynamic allocation of non-garbage-collected memory, pointers to non-heap data, and pointer arithmetic. The CLI distinguishes explicitly between *verifiable* code, which cannot use these features, and *unverifiable* code, which can. (Verifiable code must also follow a host of other rules.)

Miscellaneous Again for the sake of multiple languages, the CLI supports global data and functions, local variables whose shapes and sizes are not statically known, optional detection of arithmetic overflow, and rich facilities for "scoped" security and access control.

As in the JVM, every CLI thread has a small set of base registers and a stack of method call frames, each of which contains an array of local variables and an operand stack for expression evaluation. Each frame also contains a local memory pool for variables of dynamic and elaboration-time shape. Incoming parameters have their own separate space in the CLI; in the JVM they occupy the first few slots of the local variable array.

The Common Type System

The VES and CIL provide instructions to manipulate data of certain built-in types. A few additional types are predefined, and have built-in names in CLI metadata. To these, the CTS adds a wide variety of type constructors. For each, it defines both behavior *and* representation. No single language provides all the types of the CTS, but (with occasional compromises) each provides a subset.

The Common Language Specification (CLS) defines a subset of the CTS intended for cross-language interaction. It omits several type constructors provided by the CTS, and places restrictions on others. Standard libraries (collection classes, XML, network support, reflection, extended numerics) restrict themselves (with occasional exceptions) to types in the CLS. Not all languages support the full CLS; code written in those languages cannot make use of library facilities that require unsupported types.

Built-in Types The VES and CIL provide instructions to manipulate the following types:

- Integers in 8-, 16-, 32-, and 64-bit lengths, both signed and unsigned
- "Native" integers, of the length supported by the underlying hardware, again both signed and unsigned

- IEEE floating point, both single and double precision
- Object references and "managed" pointers

Managed pointers are different from references: while typed, they don't necessarily point to the beginning of a dynamically created object. Specifically, they can refer to fields within an object or to data outside the heap. The CIL makes sure these pointers are known to the garbage collector, which must avoid reclaiming any object *O* when a managed pointer refers to a field inside *O*. More details on pointers and references can be found in Sidebar C-16.8.

Beyond the basic hardware-level types, CLI metadata treats Booleans, characters, and strings as built-ins. Booleans and characters are manipulated in the VES using instructions intended for short integers; strings are manipulated by accessing their internal structure.

Constructed Types To the built-in types, the CTS adds the following:

Dynamically allocated instances of class, interface, array, and delegate types. These are the things to which references (the built-in type) can refer. Arrays can be multidimensional, and are stored in row-major order. Delegates are closures (subroutine references paired with referencing environments).

Methods — function types.

Properties — getters and setters for objects.

Events — lists of delegates, associated with an object, that should be called in response to changes to the object.

Value types — records (structures), unions, and enumerations.

Boxed value types — values embedded in a dynamically allocated object so that one can create references to them.

Function pointers — references to static functions: type-safe, but without a referencing environment.

Typed references — pointers bundled together with a type descriptor, used for C-style variable argument lists.

Unmanaged pointers — as in C, these can point to just about anything, and support pointer arithmetic. They *cannot* point to garbage-collectible objects (or parts of objects) in the heap.

With these type constructors come extensive semantic rules, covering such topics as identity and equality,[1] casting and coercion, scoping and visibility, mix-in inheritance, hiding and overriding of members, memory layout, initialization, type safety, and verification. The details occupy hundreds of pages in the CLI documentation.

1 These are reminiscent of the relationships tested by `eq` and `eqv` in Scheme, as discussed in Sections 7.4 and 11.3.3.

The Common Language Specification Because no single language implements the entire CTS, one cannot use arbitrary CTS types in a general-purpose interface intended for use from many different languages. The Common Language Specification (CLS) defines a subset of the CTS that most (though not all) languages can accommodate. Among other things, it omits several of the types provided by the CTS, including signed 8-bit integers; unsigned native, 16-, 32-, and 64-bit integers; boxed value types; global static fields and methods; unmanaged pointers; typed references; and methods with variable numbers and types of arguments. The CLS also imposes a variety of restrictions on the use of other types. It establishes naming conventions, limits the use of overloading, and defines the operators and conversions that programs can assume are supported on built-in types. It requires a lower bound of zero on each dimension of array indexing. It prohibits fields and static methods in interfaces. It insists that a constructor be called exactly once for each created object, and that each constructor begin with a call to a constructor of its base class. None of these restrictions applies to program components that operate only within a given language.

Generics As described in Section C-7.3.2, generics were added to Java and C# in very different ways. Partly to avoid the need to modify the JVM, Java generics

DESIGN & IMPLEMENTATION

16.8 References and pointers

The reference and pointer types of the CTS are a source of potential confusion. In a language like Java, reference types provide the only means of indirection. They refer to dynamically allocated instances of class, interface, and array types. Managed pointers provide additional functionality for languages like C# and Microsoft's C++/CLI (formerly Managed C++), which permit references to the insides of objects and to values outside the CLI heap. Managed pointers are understood by the garbage collector, and can be used in type-safe code: If a managed pointer *p* refers to a field of object *O*, then the collector will know that *O* is live. It will also update *p* automatically whenever it moves *O*.

Unmanaged pointers exist for the sake of languages like C. They are incompatible with garbage collection, and cannot point to objects in the heap. They are also incompatible with type safety, and cannot be used in verifiable code.

Typed references (`typedrefs`) in the CLI include the information needed to correctly manipulate references to values (e.g., in variable argument lists) whose type cannot be statically determined.

Version 2.0 of the CLI introduced *controlled-mutability* managed pointers (also known, somewhat inaccurately, as *read-only* pointers). Operations on these pointers are constrained to prevent modification of the referenced object. Read-only pointers are used in boxing and array contexts where generics require the ability to generate a pointer to data of a value type, but modification of that data might not be safe.

were defined in terms of *type erasure*, which effectively converts all generic types to `Object` before generating bytecode. C# generics were defined in terms of *reification*, which creates a new concrete type every time a generic is instantiated with different arguments. Reified generics have been supported directly by the CLI since .NET version 2.0, introduced by Microsoft in 2005 and codified by ECMA and ISO in 2006.

EXAMPLE 16.39 Generics in the CLI and JVM

Reified generic types are fully described in CLI metadata, allowing full type checking and reflection. Consider the following code in C#:

```
class Node<T> {
    public T val;
    public Node<T> next;
}
...
Node<int> n = new Node<int>();
Console.WriteLine(n.GetType().ToString());
```

If `Node` is an outermost class, the final line will print ``Node`1[System.Int32]``. The equivalent code in Java (running on the JVM) will simply print `class Node`. To support generics, CLI version 2 extended the rules for type compatibility and verification, and introduced new versions of several CIL instructions. ■

Metadata and Assemblies

Portable Executable (PE) *assemblies* are the rough equivalent of Java `.jar` files: they contain the code for a collection of CLI classes. PE is based on the Common Object File Format (COFF), originally developed for AT&T's System V Unix. It is the native object file format for Windows and DOS systems, extended to accommodate CIL as an optional instruction set. Given the requirements of native-code executable files (e.g., relocation—see Section 15.4), PE is quite a bit more complicated than Java `.class` and `.jar` format. A PE assembly contains a general-purpose PE header, a special CLI header, metadata describing the assembly's types and methods, and CIL code for the methods.

The metadata of an assembly has a complex internal structure. (A diagram of the interconnections among some two dozen different kinds of tables fills two pages of the annotated CLI standard [MR04, pp. 322–323].) The metadata begins with a *manifest* that specifies the files included and directly referenced, the types exported and imported, versioning information, and security permissions. This is followed by descriptions of all the types, and signatures for all the methods. Unlike the Java constant pool, the metadata of an assembly is not directly visible to the assembly's code; it may be rearranged by the JIT compiler in implementation-dependent ways, so long as it remains available to reflection routines at run time (obviously, those routines are also implementation dependent).

The Common Intermediate Language

Just as the CLI VES bears a strong resemblance to the JVM, CIL bears a strong resemblance to Java bytecode. Version 6 of the ECMA standard defines some

219 instructions, most with single-byte opcodes. Most instructions take their arguments from, and return results to, the operand stack of the current method frame. Others take explicit arguments representing variables, types, or methods. Java bytecode and CIL are similarly dense—they require roughly the same number of bytes per instruction on average.

Many of the differences between the two intermediate languages are essentially trivial. Java bytecode is big-endian; CIL is little-endian. Java bytecode has explicit instructions for monitor entry and exit; these are method calls in the CLI. CIL allows arbitrary offsets for branches; Java bytecode limits them to 64K bytes.

A few more significant differences stem from the assumption that CIL will always be JIT-compiled, as described in Sidebar C-16.7. The most obvious difference here is that Java bytecode encodes type information explicitly in opcodes, while CIL requires it to be inferred from arguments. CIL also includes an explicit instruction (`ldtoken`) that will push a "run-time handle" for a method, type, or field. While the metadata of a CIL assembly must all be available at run time, its format may be implementation dependent; the JIT compiler translates `ldtoken` into machine code consistent with that format. In the JVM, the class file constant pool is assumed to be available at run time, in its standard format; an ordinary "load constant" instruction suffices to push the desired reference.

A more subtle difference is the separation of arguments from local variables in the CLI (they share one array in the JVM). Separate arrays admit special one-byte load instructions for both the first few arguments and the first local variables, without requiring that they have interleaved slots; this in turn may make it easier to generate object code in which arguments occupy contiguous locations in memory (as, for example, in the argument build area of the stack described in Section C-9.2.2).

Finally, as already suggested, several features of CIL, not found in Java bytecode, stem from the need to support multiple source languages. We have noted that the CLI provides value types, reference parameters, and optional overflow checking on arithmetic; all of these are reflected in the CIL instruction set. There are also several extra ways to make subroutine calls. Where Java bytecode supports only static, virtual, and dynamic method invocations, CIL has (1) nonvirtual method calls, as in C++ (these implicitly pass `this`, as virtual calls do); (2) indirect calls (i.e., calls through function pointers); (3) tail calls, which discard the caller's frame; and (4) *jumps*, which redirect control to a method after executing some optional prologue (e.g., for `this` pointer adjustment in languages with multiple inheritance; see Section C-10.6).

EXAMPLE 16.40
CIL for a list insert operation

To illustrate CIL, let us return to the linked-list set of Example 16.3. The declarations given there are valid in both Java and C#. The `insert` method for this class appears in Figure C-16.7. C# source (which is again identical to the Java version) is on the left; a symbolic representation of the corresponding CIL is on the right. As in Example 16.3, there are many examples of special one-byte load and store instructions (here specified with a *.index* suffix on the opcode), and of instructions that operate implicitly on the operand stack. ■

```
public void insert(int v) {
    node n = head;
    while (n.next != null
           && n.next.val < v) {
        n = n.next;
    }
    if (n.next == null
        || n.next.val > v) {
        node t = new node();
        t.val = v;
        t.next = n.next;
        n.next = t;
    } // else v already in set
}
```

```
.method private hidebysig
       instance default void insert (int32 v)  cil managed
{
    // Method begins at RVA 0x210c      // RVA == relative
    // Code size 108 (0x6c)             //     virtual address
    .maxstack 3
    .locals init (
            class LLset/node    V_0,    // n
            class LLset/node    V_1)    // t
    IL_0000:  ldarg.0
    IL_0001:  ldfld class LLset/node LLset::head
    IL_0006:  stloc.0
    IL_0007:  br IL_0013        // jump to header of rotated loop
    IL_000c:  ldloc.0           // n -- beginning of loop body
    IL_000d:  ldfld class LLset/node LLset/node::next
    IL_0012:  stloc.0           // n = n.next
    IL_0013:  ldloc.0           // n -- beginning of loop test
    IL_0014:  ldfld class LLset/node LLset/node::next
    IL_0019:  brfalse IL_002f   // exit loop if n null
    IL_001e:  ldloc.0           // n
    IL_001f:  ldfld class LLset/node LLset/node::next
    IL_0024:  ldfld int32 LLset/node::val
    IL_0029:  ldarg.1           // v
    IL_002a:  blt IL_000c       // continue loop
    IL_002f:  ldloc.0           // n
    IL_0030:  ldfld class LLset/node LLset/node::next
    IL_0035:  brfalse IL_004b
    IL_003a:  ldloc.0           // n
    IL_003b:  ldfld class LLset/node LLset/node::next
    IL_0040:  ldfld int32 LLset/node::val
    IL_0045:  ldarg.1           // v
    IL_0046:  ble IL_006b
    IL_004b:  newobj instance void class LLset/node::'.ctor'()
    IL_0050:  stloc.1           // t
    IL_0051:  ldloc.1           // t
    IL_0052:  ldarg.1           // v
    IL_0053:  stfld int32 LLset/node::val
    IL_0058:  ldloc.1           // t
    IL_0059:  ldloc.0           // n
    IL_005a:  ldfld class LLset/node LLset/node::next
    IL_005f:  stfld class LLset/node LLset/node::next
    IL_0064:  ldloc.0           // n
    IL_0065:  ldloc.1           // t
    IL_0066:  stfld class LLset/node LLset/node::next
    IL_006b:  ret
} // end of method LLset::insert
```

Figure 16.7 **C# source and CIL for a list insertion method.** Output on the right was produced by the Mono project's `mcs` (compiler) and `monodis` (disassembler) tools, with additional comments inserted by hand. Note that the compiler has rotated the test to the bottom of the `while` loop, which occupies lines `IL_000c` through `IL_002a` in the output code.

Verification As we have noted, the CLI distinguishes between *verifiable* and *unverifiable* code. Verifiable code must satisfy a large variety of constraints that guarantee type safety and catch many common programming errors. In particular, the VES can be sure that a verifiable program will never access data outside its logical address space. Among other things, this guarantee ensures fault containment for verifiable modules that share a single physical address space.

Unverifiable code can make use of unsafe language features (e.g., unions and pointer arithmetic in C), but must still conform to more basic rules for validity (well-formedness) of CIL. Together, the components of the VES (i.e., the JIT compiler, loader, and run-time libraries) *validate* all loaded assemblies, and *verify* those that claim to be verifiable. Any standard-conforming implementation of the CLI must run all verifiable programs. Optionally, it may also run validated but not verifiable programs.

As in the JVM, verification requires data flow analysis to check type consistency and lack of underflow and overflow in the operand stack. The CLI standard requires verifiable routines to specify that all local variables are initialized to zero. CLI implementations typically perform definite assignment data flow analysis anyway, to identify cases in which those initializations can safely be omitted. The standard also requires numerous checks on individual instructions. Many of these are also performed by the JVM. Local variable references, for example, are statically checked to make sure they lie within the declared bounds of the stack frame. Other checks stem from the presence of unsafe features in the CLI. Verifiable code cannot use unmanaged pointers or unions, for example, nor can it perform most indirect method calls.

✓ CHECK YOUR UNDERSTANDING

38. Summarize the architecture of the Common Language Infrastructure. Contrast it with the JVM. Highlight those features intended to facilitate cross-language interoperability.

39. Describe how the choice of just-in-time compilation (and the rejection of interpretation) influenced the structure of the CLI.

40. Describe several different kinds of references supported by the CLI. Why are there so many?

41. What is the purpose of the Common Language Specification? Why is it only a subset of the Common Type System?

42. Describe the CLI's support for *unsafe* code. How can this support be reconciled with the need for safety in embedded settings?

16 Run-Time Program Management

16.5 Exercises

16.14 Using your local implementations of Java and C#, compile the code of Figures 16.2 and C-16.7 all the way to machine language. Disassemble and compare the results. Can all the differences be attributed to variations in the quality of the compilers, or are any reflective of more fundamental differences between the source languages or virtual machines?

16.15 Rewrite the list insertion method of Example C-16.40 in F# instead of C#. Compile to CIL and compare to the right side of Figure C-16.7. Discuss any differences you find.

16.16 Building on the previous exercise, rewrite your list insertion routine (both C# and F# versions) to be generic in the type of the list elements. Compare the generic and nongeneric versions of the resulting CIL and discuss the differences.

16.17 Extend your F# code from Exercise C-16.16 to include list removal and search routines. After finding and reading appropriate documentation, package these routines in a library that can be called in a natural way not only from F# but also from C#.

16 Run-Time Program Management

16.6 Explorations

16.24 Learn the details of the CLI verification algorithm (Partition III, Section 1.8 of the ECMA standard, version 4 [Int12a]). Pay particular attention to the rules for *merging* compatible types at joins in the control flow graph, and for dealing with generics.

16.25 Learn more about the .NET Language-Integrated Query mechanism (LINQ), mentioned in Example 16.29. Discuss its use of attributes. Write a program that uses it to interface to a database through SQL. Write another program that uses it to process the elements of a set from the `System.Collections` library.