Computability

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1 Introduction

In the last two chapters, we have introduced several important computational models, including Turing machines, and Chomsky's hierarchy of formal grammars. In this chapter, we will explore the limits of mechanical computation as defined by these models. We begin with a list of fundamental problems for which automatic computational solution would be very useful. One of these is the *universal simulation problem*: can one design a single algorithm that is capable of simulating *any* algorithm? Turing's demonstration that the answer is *yes* [Turing, 1936] supplied the proof for Babbage's dream of a single machine that could be programmed to carry out any computational task. We introduce a simple Turing machine programming language called "GOTO" in order to facilitate our own design of a universal machine. Next, we describe the schemes of *primitive recursion* and *μ-recursion*, which enable a concise, mathematical description of computable functions that is independent of any machine model. We show that the μ-recursive functions are the same as those computable on a Turing machine, and describe some computable functions including one that solves a second problem on our list.

The success in solving ends there, however. We show in the last section of this chapter that all of the remaining problems on our list are *unsolvable* by Turing machines, and subject to the *Church-Turing thesis*, have no mechanical or human solver at all. That is to say, there is no Turing machine or physical device, no stand-alone product of human invention, that is capable of giving the correct answer to all—or even most—instances of these problems. The implication we draw is that in order to solve *some* important instances of these problems, human ingenuity is needed to guide powerful computers down the paths felt most likely to yield the answers. To cite Raymond Smullyan quoting P. Rosenbloom, the results on unsolvability imply that "man can never eliminate the necessity of using his own cleverness, no matter how cleverly he tries."

Almost the first consequence of formalizing computation was that we can formally establish its limits. Kurt Gödel showed that the process of proof in any formal axiomatic system of logic can be simulated by the basic arithmetical functions that computation is made of. Then he proved that any sound formal system that is capable of stating the grade-school rules of arithmetic can make statements that can be neither be proved nor disproved in the system. Put another way, every sound formal system is *incomplete* in the sense that there are mathematical truths that cannot be proved in the system. Turing realized that Gödel's basic method could be applied to computational models themselves, and thus proved the first computational unsolvability results. Since then, problems from many areas, including group theory, number theory, combinatorics, set theory, logic, cellular automata, dynamical systems, topology, and knot theory, have been shown to be unsolvable. In fact, proving unsolvability is now an accepted "solution" to a problem. It is just a way of saying that the problem is too general for a computer to handle—that supplementary information is needed to enable a mechanical solution.

Since Turing machines capture the power of mechanical computability, our study will be based on Turing's model. In the next section, we describe a Turing machine as a computer that can run programs written in a very simple language we call the "GOTO Language." This formalism is equivalent to Chapter 30's description of Turing machines using the standard 7-tuple notation. Our language provides an alternate way to write programs and makes proofs about Turing machines more intuitive.

2 Computability and a Universal Program

Turing's notion of mechanical computation was based on identifying the *basic steps* in any mechanical computation. He reasoned that an operation such as numerical multiplication is not primitive, because it can be divided into simpler steps such as using the times-table on individual pairs of digits, shifting, and adding. Addition itself can be broken down into simpler steps such as adding the lowest digits, computing the carry, and moving to the next digit. Turing concluded that the most basic features of mechanical computation are the ability to read and write on a storage medium, the ability to move about on that medium, and the ability to make simple logical decisions. Turing chose the storage medium to be a single linear *tape* divided into cells. He showed that such a tape could model spatial memory in three (or any number of) dimensions through the use of indexed co-ordinates. With much care he argued that human sensory input could be encoded by strings over a finite alphabet of cell symbols called the *tape alphabet*. (This bold discretization of sensory experience now seems a harbinger of the digital revolution that was to follow.) A decision step enables the computer to exert local control over the sequence of actions. Turing restricted the next action performed to be in a cell neighboring the one on which the current action occurred, and showed how non-local actions can be simulated by successions of steps of this kind. He also introduced an instruction to tell the computer to stop.

In summary, Turing proposed a model to characterize mechanical computation as being carried out as a sequence of instructions. Our "GOTO" formalism provides the following five kinds of instructions. Here i stands for a tape symbol and j stands for a line number.

PRINT i
MOVE RIGHT
MOVE LEFT
IF i IS SCANNED GOTO LINE j
STOP

When we speak about programs recognizing languages rather than computing functions, we replace STOP by statements ACCEPT and REJECT, each of which need occur only once in a program.

A program in this language is a sequence of instructions or "lines" numbered 1 to k. The input to the program is a string over a designated *input alphabet* Σ, which we take to be $\{0, 1\}$ throughout this chapter. The tape alphabet includes Σ and a special *blank* character B representing an empty cell, and may (but need not) contain other symbols. The input is stored on the tape, with the read head scanning the first symbol (or B if the input is empty), before the computation begins.

How much memory should we allow the computer to use? Rather than postulate that the tape is actually infinite—an unrealistic assumption—we prefer here to say that the tape has expandable boundaries. Initially the input defines the two boundaries of the tape. Whenever the machine moves left of the left boundary or right of the right boundary, a new memory cell containing the blank is attached. This convention clarifies what we mean by saying that if and when the machine halts by reaching the STOP instruction, the "result" of the computation is the entire content of the tape.

We present an example program written in the GOTO language. This program accomplishes the simple task of doubling the number of initial 1s in the input string. Informally, the program achieves its goal as follows: When it reads a 1, it changes the 1 to a 0, moves left looking for a new cell, and writes a 1 in that cell. Then it returns rightward to the 0 that marks where it had been, rewrites it as a 1, and moves right to look for more 1s. If it immediately finds another 1 it repeats the process from line 1, while if it doesn't, it halts right there. This program even has a "bug"—it "should" leave strings that do not begin with a 1 unchanged, but instead it alters them.

The main change from the traditional Turing machine formalism of Chapter 30 is that we have replaced "states" by line numbers. A Turing machine of the former kind can always be simulated in our GOTO language by making blocks of successive lines, themselves divided into sub-blocks for each character that are headed by "IF" statements, carry out the instructions for each state. Our formalism makes many programs more succinct and closer to programmers' experience, and highlights the role of (conditional) GOTO instructions in setting up loops and enabling statements to be repeated. Despite the popular scorn of `goto` statements, this feature is ultimately the most important aspect of programming and can be found in every imperative-style programming language—at least in the code produced by the compiler if the language has no `goto` instruction itself. Indeed, the above example could be rendered into a structured programming language such as C as follows,

```
do {
    do { PRINT 0; MOVE LEFT; } while (1 is scanned);
```

```
1 PRINT 0
2 MOVE LEFT
3 IF 1 IS SCANNED GOTO LINE 2
4 PRINT 1
5 MOVE RIGHT
6 IF 1 IS SCANNED GOTO LINE 5
7 PRINT 1
8 MOVE RIGHT
9 IF 1 IS SCANNED GOTO LINE 1
10 STOP
```

Figure 1: The doubling program in the GOTO language.

```
        PRINT 1;
        do MOVE RIGHT; while (1 is scanned);
        PRINT 1; MOVE RIGHT; }
    while (1 is scanned);
```

and a C compiler (using a character array `tape[i]` and `++i`, `--i` for the moves) might plausibly convert this into something exactly like our GOTO program!

The simplicity of the GOTO language is rather deceptive. As the above example hints, any program in any known high-level programming language can be converted into an equivalent GOTO program, under suitable conventions on how inputs and outputs are represented on the tape. (If the program only reads from the standard input stream and writes to the standard output stream, then no such conventions are necessary.) There is strong reason to believe that any mechanical computation of any future kind can be expressed by a suitable GOTO program. Note, however, that a program written in the GOTO language need not always halt; i.e., on certain inputs the program may never reach a STOP instruction. On such inputs we say that the output of the program is *undefined*.

Now we can give a precise definition of what we mean by an *algorithm*, attempting to rule out this last situation. An algorithm is *any program written in the* GOTO *language that has the additional property of halting on all inputs.* Such programs will be called *halting programs*, and correspond to "total" deterministic Turing machines in Chapter 30. When we consider *decision problems*, which have yes/no answers, halting programs are required to end their computation with either an ACCEPT or a REJECT statement, on any input.

2.1 Some Computational Problems

We begin by listing a collection of computational problems for which a mechanical solution can be very helpful. By a mechanical solution, we mean a step-by-step process that takes into account all possible inputs, and that can be executed without any human assistance once a certain input is provided. An algorithm is required to work correctly on all instances.

We now list some problems that are fundamental either because they are inherently important or because they played a historical role in the development of computation theory. For the first four, P stands for a program in our GOTO language, and x is a string over the input alphabet, which we fix to be $\{0, 1\}$.

1. *Universal simulation.* Given a program P and an input x to P, determine the output (if any) that P would produce on input x.

2. *Halting problem.* Given P and x, output 1 (for yes) if P would halt when given input x, and 0 (for no) if P would not halt.

3. *Type-0 grammar membership.* Given a type-0 grammar G and a string x, determine whether x can be derived from the start symbol of G.

4. *String compression.* Given a string x, find the shortest program P such that when P is started with empty tape, P eventually halts with x as its output. Here "shortest" means that the total number of symbols in the program's instructions is as small as possible.

5. *Tiling.* Given a finite set T of tile types, where all tiles of a type are unit squares with the same four colors on their four edges, determine whether every finite rectangle can be *tiled* by T. If k and n are the integer sides of the rectangle, being tiled means that kn tiles drawn

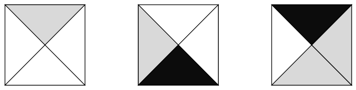
from T can be arranged so that every two tiles that share an edge have the same color at that edge.

6. *Linear programming.* Given some number k of linear inequalities in n unknowns, determine whether there is an assignment of n values to the unknowns that satisfies all the inequalities.

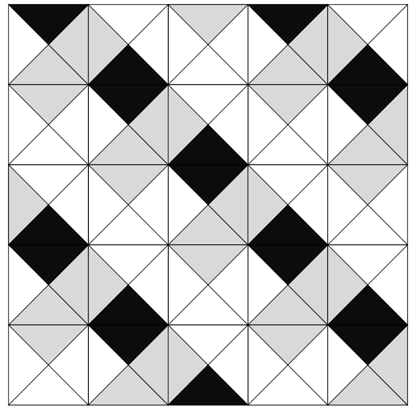
7. *Integer equations.* Given k-many polynomial equations in n unknowns, determine whether there is an assignment of n integers to the unknowns that satisfies all the equations.

Some remarks about the above problems: A solution to Problem 1 realizes Babbage's objective of a single program or machine capable of simulating all programs P. For cases where P run on x would never halt and produce output, we have left open whether we require the solution itself to halt and detect this fact—i.e., to be an algorithm. For any such algorithm to exist, there must be an algorithm to solve Problem 2, which is a yes/no *decision problem*. An algorithm for Problem 2 would be a boon to reliable software design, since it could be used to test whether a given block of code can cause infinite loops. Problem 3 is another decision problem; its solution would be useful for natural-language processing and much more. Problem 4 is a function-computation problem of central importance in information theory. For illustration, think of x as a large amount of scientific data for which we seek a concise theory P that can generate and hence explain it. A famous example is Kepler's laws, which explained Tycho Brahe's voluminous and meticulous observational data. Problem 4 thus asks whether the heart of science (to paraphrase Occam's Razor, "finding the simplest explanation that fits the facts") can be done automatically on a computer. The tiles in Problem 5, which we introduced in detail in Chapter 30, are sometimes named after Hao Wang [Wang, 1961], who wrote the first research paper about them.

Figure 2(a) shows an example of a set T of tile types, and Figure 2(b) shows how tiles drawn from T can be used to tile a 5×5 square area. The tiling problem is not merely an interesting puzzle. It has been is an art form pursued by artists from many cultures for centuries. Tiling problems have deep significance in combinatorics, algebra and formal languages. Note that our decision problem does not ask simply whether a given $k \times n$ rectangle can be tiled, but whether—given T—*all* $k \times n$ rectangles can be tiled via T. The full problem of linear programming adds to Problem 6 a clause saying: if there exist *feasible solutions*, i.e. assignments that satisfy all the so-called linear *constraints*, find one that maximizes (or minimizes) a given *objective function* (or



(a)



(b)

Figure 2: An example of tiling.

cost function). This problem has central importance in economics, game theory, and operations research. Problem 7 is called *Hilbert's tenth problem*, and was one of twenty-four that David Hilbert posed as challenges for the new century at the International Congress of Mathematicians in 1900. It goes back two thousand years to the mathematician Diophantus' study of these so-called *Diophantine equations*. Actually, Hilbert posed the "meta-problem" of finding an algorithm that can solve any Diophantine equation, or at least tell whether it has a solution.

Recall from Chapter 30 that a decision problem is *decidable* if it has an algorithm, and *undecidable* otherwise. If our program correctly evaluates all instances for which the answer is "yes," but may fail to halt on some instances for which the answer is "no," then the program is a *partial decision procedure*, and the problem is *partially decidable*. A partially decidable problem, however, is undecidable—unless you can find an algorithm that removes the word "partially." Likewise, if our program correctly outputs $f(x)$ whenever $f(x)$ is defined, but may fail to halt when $f(x)$ is undefined, then the partial function f is *partial computable*.

In the remainder of the subsection, we present some simple algorithm design techniques and sketch how they make progress on solving some of these problems and special cases of them. These techniques may seem too obvious to warrant explicit description. However, we feel that such a description will help new readers to appreciate the limits on information processing that make certain problems undecidable.

2.1.1 Table Look-up

For certain functions g it can be advantageous to create a table with one column for inputs x and one for values $g(x)$, looking up the value in the table whenever an evaluation $g(x)$ is needed. A function f that is defined on an infinite set such as Σ^* cannot have its values enumerated in a finite table in this manner, but sometimes the infinite table for f can be described in a finite way that constitutes an algorithm for f. Moreover, tables for other functions g may help the task of computing f, such as the digit-by-digit times-table used in multiplying integers of arbitrary size. These ideas come into play next.

2.1.2 Bounding the Search Domain

Many solutions to decision problems involve finding a *witness* that proves a "yes" or "no" answer for a given instance. The term reflects an analogy to a criminal trial where a key witness may determine the guilt or innocence of the defendant. Thus the first step in solving many decision problems is to identify the right kind of witness to look for. For example, consider the problem of determining whether a given number N is prime. Here a (counter-) witness would be a factor of N (other than 1 and N itself). If N is composite, it is easy to prove by simple division that the witness' claim is correct.

In cases where the given number is prime, a witness of a different kind needs to be searched for. This search may involve integers larger than N, and trying to summon every integer sequentially as a witness would violate the requirement of an algorithm to terminate in finite number of steps. *This is often the main challenge in establishing decidability.* The difficulty can be surmounted if, based on the structure of the problem, we can establish ahead of time an upper bound such that if any witness exists at all, one exists that meets the bound. Then a sufficient body of potential witnesses can be examined in a finite number of steps. In the case of composite N, the bound is N itself. For prime N, there is a known polynomial p such that a witness exists in the numbers between 1 and $2^{p(n)}$, where n is the number of digits in N, according to a certain witnessing scheme whose test for correct claims is easy to compute. This kind of "polynomial size-bounded witnessing scheme" characterizes the important complexity class NP, and is discussed much further in Chapter 33.

For another example, let us consider the special case of Problem 3 where the given G is a

Type-1 grammar, and we wish to determine whether a given string x can be generated from the start symbol S of G. A witness in this case can be a sequence of sentential forms starting from S and ending with x that forms a valid derivation in G. The length of x imposes a limit on the size of such sentential forms because G has no length-decreasing productions, and this in turn defines a (much larger) limit on the number of sequences that need be considered before all possibilities are exhausted. Readers may find the details in a standard text such as [Hopcroft and Ullman, 1979].

For one more example, consider the full version of the linear programming problem where one wishes to maximize a linear objective function f over the set of feasible solutions s. This set may be infinite, and so a table-lookup through all values $f(s)$ cannot be used. However, it is possible to reduce the search domain to a finite set as follows. The feasible solutions form a collection n-dimensional space, where n is the number of variables plus the number of constraints) known as a *convex polytope*. Unless the polytope is empty or unbounded—cases that can be detected and resolved—the polytope has a finite number of "corner" points, which are similar to the vertices of a polygon, and which are easily computed. In this case, it is known that a linear objective function attains its maximum value at one (or more) of these corner points. Thus we know the problem is decidable via table-lookup of values at the corner points. In practice, there are intelligent algorithms that find a maximum-giving corner point after searching (usually) only a small part of this table.

2.1.3 Use of Subroutines

This is more a programming technique than an algorithm design tool. The idea is to use one program P as a single step in another program Q. Building programs from simpler programs is a natural way to deal with the complexity of the programmer's task. A simple example is using a lookup to the times table as a subroutine in multiplying two integers i and j. Let us examine this in the context of designing Turing machines, where i and j are represented on the tape by the string 1^i01^j (namely, i 1s followed by a 0 and then by j 1s). The basic idea of our GOTO program is to duplicate the string of i 1s $j-1$ times, meanwhile erasing the string 1^j bit-by-bit to count the iterations. A little thought reveals that our earlier GOTO program in Figure 2 can *almost* be used verbatim as a subroutine to call $j-1$ times. The only hitch is that the first call would run $2i$-many 1s together so that further calls would duplicate too many 1s. To fix the problem, we introduce a new tape symbol 2, using two initial steps to convert the tape to 21^i01^j, and "patch" the subroutine so that it will not overwrite this 2. The new subroutine can be called by a line "k: IF 2 IS SCANNED GOTO m," where m is the number of the first line in the subroutine, and can return control to the point of call by replacing its STOP instruction by "IF 0 IS SCANNED GOTO $k+1$." Careful writing will ensure that this latter 0 is the one initially separating 1^i from 1^j. The remaining details are left to the interested reader, while performing a similar patch without using a new symbol "2" is left to the obsessive reader. This subroutine mechanism is in fact no different from the one programmers in BASIC have used for decades.

2.2 A Universal Program

We will now solve Problem 1 by arguing the existence of a program U written in the GOTO language that takes as input a program P (also written in the GOTO language) and data x for P, and that produces the same output as P does on input x, *if* $P(x)$ halts and produces output at all. The last caveat hints that we shall only achieve a partial solution, formally showing only that the function $U(P,x) = P(x)$ is partial computable.

For convenience, we assume that all programs written in the GOTO language use the fixed alphabet $\{0,1,B\}$. Since we have thus far used the full English alphabet for the notation of our GOTO programs, we must first address the issue of what the formal input to the program U will look like. This problem can be circumvented by *encoding* each instruction using only 0 and 1. The idea of such an encoding should not be mysterious—we could refer to the 0-1 encoding defined by the ASCII standard, which the terminal used to type this chapter has already carried out for these example programs. However, we prefer the more-succinct encoding defined by Table 1.

To encode an entire program, we simply write down in order (without the line numbers) the code for each instruction as given in the table. For example, here is the code for the doubling program shown in Figure 2:

$$0001001011110110001101001111111011000110100111011100.$$

Note that the encoded string preserves all the information about the program, so that one can easily reverse the process to decode the string into a GOTO program. From now on, if P is a program in the GOTO language, *code*(P) will denote its binary encoding. When there is no confusion, we will

Table 1: Encoding the GOTO instructions.

Instruction	Code
PRINT i	0001^{i+1}
MOVE LEFT	001
MOVE RIGHT	010
IF i IS SCANNED GOTO j	$0111^{j}01^{i+1}$
STOP	100

identify P and $code(P)$. We may also assume that all programs P have a unique STOP instruction that comes last. This convention ensures that a input string to U of the form $w = code(P)x$ can be parsed into its P and x components. (When we consider decision problems we will use the code 100 for a unique final ACCEPT instruction, and assign some other code to REJECT.) Before proceeding further, readers may test their understanding of the encoding/decoding process by decoding the following string: 0100111011001001.

The basic idea behind the construction of a universal program is simple, although the details involved in actually constructing one are substantial. Turing in his original paper [Turing, 1936] exhibited a universal program in glorious gory detail, while simpler constructions may be found in more-recent sources such as [Robinson, 1991]. Here we will content ourselves with a sketch that conveys the central ideas.

U has as its input a string w of the form $code(P)x$. (If U is given an input string not of this form, it can detect the flaw and immediately stop.) To simulate the computational steps of P on input x, U divides its tape into two segments, one containing the program P, and one modeling the contents of the tape of P as it changes with successive moves. The computation by U consists of a sequence of *cycles*, each of which simulates one step by P and is analogous to an *REW cycle* (for "read-evaluate-write") in many real computer systems.

To execute a cycle, U first needs to know the cell that the "virtual" tape head of P is currently scanning, and the instruction P is currently executing. We can assist U by extending its own work alphabet to include new "alias" symbols $0', 1', B'$ for the characters of P. U maintains the condition that there is exactly one aliased symbol in the "P" segment of its tape that marks the encoding of the current instruction, and exactly one in the other segment that marks the cell currently scanned by P. For example, suppose that after thirty-nine steps, P is reading the fourth symbol from the left on its tape containing 01001001. Then the second tape segment of U after thirty-nine cycles consists of the string $0100'1001$. We can further assist U by adding a symbol $\wedge$ to divide the two segments, although the unique STOP instruction itself could serve as the divider. The computation by U on an input $w = \mathit{code}(P)x$ can begin with some steps that prime the first symbol of $\mathit{code}(P)$, insert a $\wedge$ before the first symbol of x (caterpillaring x one cell to the right), and prime the first symbol of x. We may suppose that each cycle by U begins with its own head scanning the $\wedge$.

At the beginning of a new cycle, U moves its head left to find the current instruction, and begins decoding it. The only information U needs to retain is which type of instruction it is, and in the case of a PRINT i or IF $i\ldots$ instruction, which character i is involved. To execute a PRINT i, MOVE RIGHT, or MOVE LEFT instruction, P unprimes the instruction, primes the next one, and marches down its tape to find the primed cell on its copy of P's tape and execute the action. It is possible that a MOVE LEFT instruction may bump into the $\wedge$, in which case U makes another call to its "caterpillar" subroutine to move P's tape over, and inserts a B' for the blank P would scan after that move. The only case that requires cumbersome action by U is an instruction IF i IS SCANNED GOTO j, when U finds that P really is scanning character i. Then U needs to find the jth instruction in the "P" part of its tape. Because we have used a unary encoding 1^j of the required line number j, it is not too difficult to write a subroutine that counts off the 1s in 1^j and advances an instruction marker each time beginning from line 1, knowing to stop when the jth instruction has been located. Finally, if the current instruction is STOP, U gleefully *erases* P, erases the $\wedge$, and unprimes the scanned symbol, leaving exactly the final output $P(x)$.

One last refinement is needed to answer the objection that U is using extra tape symbols $0', 1', B', \wedge$ that we have expressly forbidden to GOTO programs. This use can be eliminated by one more level of encoding. Give each of the seven tape symbols its own three-bit code, and make U treat blocks of three cells as single cells in the simulation that was described above. U itself can be programmed to convert its input $\mathit{code}(P)x$ to this encoding before the first cycle, and to invert it when restoring the final output $P(x)$. Then U is a bona-fide GOTO program that meets all our requirements. It is even possible to run U on input $\mathit{code}(U)w$ where $w = \mathit{code}(P)x$, producing (more slowly) the same output $P(x)$. It is important to note that the code of U itself is completely independent of any program P that might be simulated. The code of U itself is not long—a reader with good programming skill can make it shorter than the prose description we have just given.

Besides solving what was asked for in Problem 1, we have also shown that Problem 2 is partially decidable. Namely, for any "yes"-instance $w = code(P)x$ where P on input x halts, U on input w will eventually detect that fact—and the slight edit of changing U's own STOP instruction to ACCEPT will make U halt and accept w. However, on a "no"-instance where $P(x)$ does not halt, our U will blindly follow P and not halt either. The question is whether we can improve U so that it will *detect* every case in which $P(x)$ does not halt, and signal this by executing a REJECT instruction. We will see in Section 5 that all the programming skill in the world cannot produce such a U—the halting problem is *un*decidable.

Before presenting undecidability, however, we develop a fundamentally different way to formalize the notion of mechanical computability in the next section.

3 Recursive Function Theory

The main advantage of using the class of μ-*recursive* functions to define computation is their mathematical elegance. Proofs about this class can be presented in a rigorous and concise way, without long prose descriptions or complicated programs that are hard to verify. These functions need and make no reference to any computational machine model, so it is remarkable that they characterize "mechanical" computability.

An analogy to the two broad families of programming languages is in order. We have already discussed how Turing machines and our particular "GOTO" formalism abstract the essence of *imperative* programming languages, in which a program is a sequence of operational commands and the major program structures are subroutines and loops and other forms of *iteration*. By contrast, specifications in recursive function theory are *declarative*, and the major structures are forms of *recursion*. "Declarative" means that a function f is specified by a direct description of the value $f(x)$ on a general argument x, as opposed to giving steps to compute $f(x)$ on input x. Often this description is *recursive*, meaning that $f(x)$ is defined in terms of values $f(y)$ on other (usually smaller) arguments y. Programming languages built on declarative principles include Lisp, ML, and Haskell, which are known as *functional languages*. These languages have recursion syntax that is not greatly different from the recursion schemes presented here. They also draw upon Church's *lambda calculus*, which can be called the world's first general programming language. A formal proof of equivalence between lambda calculus and the Turing machine model (via a programming language called I) can be found in [Jones 1997], which presents computability theory from a programming perspective.[1]

In this section, we will describe this functional approach to computation and code some simple functions using recursion. Owing to space limitation, we will not present a complete proof that the class of μ*-recursive* functions is the same as the class of (partial) computable functions on a Turing machine. The full proof can be found in standard texts such as [Sudkamp, 1997]. All the functions we consider have one or more non-negative integers as arguments, and produce a single non-negative integer value.

Before presenting formal definitions, we qualify the above ideas with a few examples. Consider first the simple definition of a two-variable linear function by

$$h(y, z) = z + 2 * y + 1. \tag{1}$$

Here $h(y, z)$ is defined with the aid of other functions (here, plus and times) and quantities (here, 2 and 1) that presumably have already been defined or given. This is an example of an *explicit definition* because all entities on the right-hand side are known—in particular, this definition does not involve recursion. If we rewrite the infix functions $+$ and $*$ in prefix-function style as "*plus*" and "*times*," the expression becomes

$$h(y, z) = plus(z, plus(times(2, y), 1)), \tag{2}$$

and we can glimpse another hallmark of functional languages: function names can be regarded as parameters the same way that variable names can. Now consider the somewhat-similar definition of a one-variable function by

$$f(x) = f(x - 1) + 2 * (x - 1) + 1, \tag{3}$$

[1]Turing created an addendum to his seminal paper [Turing, 1936] showing that his definition of a (partial) computable function was equivalent to the one proposed by Church. The lambda calculus uses essentially a single execution scheme called *reduction* to govern its computations, and by suitable conditioning one can make this scheme carry out recursion. Another declarative language, Prolog, also fixes a single execution scheme that tries to limit the operational decisions the programmer needs to make, and also relies upon recursion.

together with a base case such as $f(0) = 0$. Here not every quantity on the right-hand side is known—one must first know $f(x-1)$ to compute $f(x)$. However, this is still "declarative" insofar as $f(x)$ is defined in terms of known quantities and values $f(y)$ for other (smaller) arguments y. The reader may check that this is a recursive definition of the squaring function.

Why use recursion? One reason is that explicit definition by itself is known not to be powerful enough to capture the essence of mechanical computation. The next two sections define the two principal schemes of recursion in recursive function theory.

3.1 Primitive Recursive Functions

The class of *primitive recursive functions* is built up from the following set of *basic functions*, which are the only ones we need to pre-suppose are "known":

1. The *successor function* S is defined for all x by $S(x) = x + 1$.
2. The *zero function* Z is defined for all x by $Z(x) = 0$. The constant 0 is also provided here.
3. For all fixed numbers i and n with $1 \leq i \leq n$, the *projection function* p_i^n is defined for all n-tuples $(x_1, x_2, \ldots, x_n)$ by $p_i^n(x_1, x_2, \ldots, x_n) = x_i$.

The primitive recursive functions are constructed from the basic functions by applications of the following two operations. The case $n = 0$ is allowed in them; a 0-variable function is the same as a constant, and a 0-tuple is the empty list.

1. *Functional composition:* Given k-many functions $g_1, \ldots, g_k$ that each take n variables, and a function h that takes k variables, one can define a function f of n variables by

$$f(x_1, \ldots, x_n = h(g_1(x_1, \ldots, x_n), g_2(x_1, \ldots, x_n), \ldots, g_k(x_1, \ldots, x_n)). \tag{4}$$

   If $g_1, \ldots, g_k$ and h are primitive recursive, then f is defined to be primitive recursive.

2. *Primitive recursion:* Given a function g that takes n variables, and a function h that takes $n + 2$ variables, one can define a function f of $n + 1$ variables by

$$f(x_1, \ldots, x_n, 0) = g(x_1, \ldots, x_n); \tag{5}$$

$$f(x_1, \ldots, x_n, S(y)) = h(x_1, \ldots, x_n, y, f(x_1, \ldots, x_n, y)). \tag{6}$$

If g and h are primitive recursive, then f is defined to be primitive recursive.

Here (5) is the *basis* and (6) is the *recursion step*. It is conventional to call $x_1, \ldots, x_n$ the *parameters* and y the *recursion variable*. From a computational viewpoint, the scheme is easy to interpret. Given integer values for variables $x_1, \ldots, x_n$ and z, how can we evaluate $f(x_1, \ldots, x_n, z)$? We start building a table T in which each row y contains the value of $f(x_1, \ldots, x_n, y)$. The basis step gives us the top row via $T[0] = f(x_1, \ldots, x_n, 0) = g(x_1, \ldots, x_n)$. Whenever we have filled a row y, we can fill the next row via the recursion step, via $T[y{+}1] = f(x_1, \ldots, x_n, S(y)) = h(x_1, \ldots, x_n, y, T[y])$. As soon as row z is filled, using y such that $z = S(y)$, we are done. The point is that provided g and h are computable, the function f is also computable. Functional composition likewise preserves computability. Moreover, since the basic functions are all total and produce non-negative values, every function that we can build up in this manner is also total and produces non-negative values.

DEFINITION 3.1 A function is said to be **primitive recursive** if it can be built up from the successor, zero and projection functions by a finite number of applications of composition and primitive recursion.

EXAMPLE 3.1 To show how the scheme of primitive recursion models the informal recursion defining the function f of one variable (so we have $n = 0$) in equation (3), take "$g()$" to be the constant 0, and take h to be the two-variable function $h(y, z) = z + 2y + 1$, which happens to be our example of "explicit definition" in (1). Then we have $f(0) = 0$ and

$$f(S(y)) = h(y, f(y)) = f(y) + 2 * y + 1.$$

With "$x - 1$" in place of "y" and "x" in place of "$S(y)$," this is the same as (3). We will return to this notational difference later.

As the prefix form (2) indicates, h itself can be built up via functional composition from the *plus* and *times* functions. It is interesting to see how the usual functions of arithmetic can themselves be constructed from the rather Spartan basis we have been given. To begin with, the constants $1, 2, \ldots$ are formally introduced by functional composition, with "$g_1()$" as the constant 0 and "h" as the successor function, via $1 = S(0)$, $2 = S(1) = S(S(0))$, $3 = S(2)$, and so on.

Example 3.2 *Addition.* Take $g(x) = x$ and $h(x, y, z) = S(z)$. Formally, g is the basis function p_1^1, and h is the functional composition of the successor function with p_3^3. Then primitive recursion gives us $plus(x, 0) = g(x) = x$ and

$$plus(x, S(y)) = h(x, y, plus(x, y)) = S(plus(x, y)) = S(x + y) = x + y + 1,$$

as we would demand. Hence this formal definition of *plus* correctly computes addition, and we may use the standard "+" notation in the formal examples that follow.

Example 3.3 *Multiplication.* Take $g(x) = 0$ and $h(x, y, z) = x + z$. Formally, g is the zero function (of one variable rather than the constant zero), and h is the functional composition of *plus* with the two functions p_1^3 and p_3^3 (so $k = 2$ here). Then primitive recursion gives us $times(x, 0) = g(x) = 0$ and

$$times(x, S(y)) = h(x, y, times(x, y)) = x + times(x, y) = x * (y + 1),$$

again as we would demand. Hence this formal definition of *times* correctly computes multiplication. Note that we had to go to some length (of making h a function of 3 variables) so that our definition exactly agrees with the formal requirements in equation (6).

Example 3.4 *Exponentiation.* Take $g(x) = 1$ and $h(x, y, z) = x * z$. Formally, g is the one-variable function that always outputs 1 and is defined by composing S and the zero function Z, while h is the same as in Example 3.3 but with *times* in place of *plus*. Then primitive recursion gives us $exp(x, 0) = g(x) = 1$ (note that even 0^0 equals 1) and

$$exp(x, S(y)) = h(x, y, exp(x, y)) = x * exp(x, y) = x^{y+1}.$$

Once again the correctness of this definition for all values of x and y is easy to verify, via a simple proof by induction that follows the recursion.

It is now straightforward to omit some of the formal apparatus and write the definitions more succinctly. For instance, the last example becomes

$$\begin{aligned} exp(x, 0) &= 1 \\ exp(x, y + 1) &= x * exp(x, y). \end{aligned}$$

This resembles a program one would actually write, especially in a language like C that does not provide exponentiation as a built-in operator.

At this point the alert reader, noting the way our schemes all involve non-negative numbers, will first wonder how on earth we can ever define *subtraction* this way. The key is that the syntax of primitive recursion allows us to define a function $P(y)$ that computes "proper subtraction by 1," and then use P to define *proper subtraction* itself. The word "proper" here means that any negative value is replaced by 0, in order to maintain our restriction to the non-negative numbers. The definitions are

$$\begin{aligned} P(0) &= 0 \\ P(S(y)) &= y \\ sub(x,0) &= x \\ sub(x,S(y)) &= P(sub(x,y)) \end{aligned}$$

For P we took $h(y,z) = y$, i.e. $h = p_1^2$, and for sub we took $h(y,z) = P(z)$. To trace this out, $sub(3,2) = P(sub(3,1)) = P(P(sub(3,0)) = P(P(3)) = P(2) = 1$, and $sub(2,3) = P(P(P(2))) = P(0) = 0$, which is the "proper" value.

Second, the reader may have felt uncomfortable defining functions in terms of "$S(y)$" rather than "y." For example, the primitive recursion for the factorial function, with 0! standardly defined to be 1, gives us

```
fact(0) = 1 | fact(y+1) = (y+1)*fact(y);
```

here "`|`" separates the base and recursion cases. This would actually be valid syntax in the programming language ML *except* that "`fact(y+1)`" is an illegal function header. The syntax of ML forces one to write it this way:

```
fact(0) = 1 | fact(y) = y*fact(y-1);
```

this is literally the example used in many texts. To make the formal equation (6) for primitive recursion reflect the syntax of programming languages, we can use P in place of S to change it to

$$f(x_1,\ldots,x_n,y) \;=\; h(x_1,\ldots,x_n,y,f(x_1,\ldots,x_n,P(y))), \tag{7}$$

or alternately make the middle argument of h be $P(y)$ instead of y. Either way, one might expect to be able to recover the function S by defining it in terms of P and the other two basis functions, just as we defined P in terms of S above. However, this is impossible—one could never define any increasing functions at all. This curious asymmetry partly explains why primitive recursion was defined the way it is. Nevertheless, if S as well as P is provided in the basis, then one can use the modified definition and obtain exactly the same class of primitive recursive functions. For instance, addition is definable by $plus(x, 0) = x \mid plus(x, y) = S(plus(x, P(y)))$, and so on. Hence primitive recursion is for the most part exactly what ML and other functional languages *do*.[2]

Finally, the reader may wonder what has become of functions defined on *strings*. A string over an alphabet Σ can always be identified with its number in the standard lexicographic enumeration of Σ^*, with ϵ corresponding to 0. Then a string function $f : \Sigma^* \to \Sigma^*$ can be called primitive recursive if the corresponding numerical function (of one variable) is primitive recursive. For instance, the function that appends a '1' to a binary string x corresponds to $2x + 2$. Cutting the other way, under some transparent encoding of negative and rational and complex numbers (etc.) by strings, one can extend the concept of primitive recursion to define addition and multiplication and nearly all familiar mathematical functions in their full generality. The meaning and proof of the following statement should now be clear; full detail can be found in [Sudkamp, 1997].

THEOREM 3.1 *Every primitive recursive function is computable by a Turing machine.*

The converse is false, however. A famous example of a computable total function that is not primitive recursive is *Ackermann's function*; this and other examples may be found in [McNaughton, 1993]. To obtain all computable functions we need to introduce one more scheme of recursion—at the inevitable cost, however, of opening a Pandora's box of functions that are no longer total.

[2] Primitive recursion has its counterpart in imperative languages as well, aside from the fact that most of them support recursion directly. The "table $T[y]$" computation above shows how primitive recursion can be simulated by a simple for-loop **for** $y = 0$ **to** z **do...end** that fixes its bounds and never alters y in the loop body. A theorem [Meyer and Ritchie] in programming languages states that the primitive recursive functions are exactly the total functions computable by programs that use only if-then-else and nested for-loops.

3.2 μ-Recursive Functions

We will add a new operation called *minimalization* that does not preserve totality. Again we restrict numerical arguments to be non-negative integers.

DEFINITION 3.2 A possibly-partial function f of n variables is defined by μ-**recursion** from a function g of $n+1$ variables, written

$$f(x_1, \ldots, x_n) = \mu y.g(x_1, \ldots, x_n, y),$$

if whenever $f(x_1, \ldots, x_n)$ is defined, it equals the least number y such that $g(x_1, \ldots, x_n, y) = 1$. If $f(x_1, \ldots, x_n)$ is undefined, there must be no y such that $g(x_1, \ldots, x_n, y) = 1$. The class of μ-**recursive functions** is the class of all functions that can be built up from the successor, zero, and projection functions by the operations of composition, primitive recursion, and μ-recursion.

The computation of $f(x_1, \ldots, x_n)$ that is implicit in Definition 3.2 can be described by building a table as before. First fill in the row $T[0] = g(x_1, \ldots, x_n, 0)$, then $T[1] = g(x_1, \ldots, x_n, 1)$, and so on. If and when one finds a y whose value $T[y]$ equals 1, halt and output y. The "if" is the big difference from the algorithm for primitive recursion, because if $g(x_1, \ldots, x_n, y)$ never takes the value 1, this procedure will never halt. This procedure is called an *unbounded search*. Compared another way to primitive recursion, μ-recursion increments its recursion variable rather than decrement it.

There is nothing special about "$= 1$" here: zero or any other constant could be used instead. Our use of 1 suggests the special case in which g is a total function that takes on only the values 0 and 1. Then we can regard its output as a Boolean truth value, with $1 =$ true and $0 =$ false, and call g a *predicate*. The class of μ-recursive functions is not changed under the restriction that g be a predicate. Then we can read the syntax "$\mu y.g(x_1, \ldots, x_n, y)$" in English as "the least y such that $g(x_1, \ldots, x_n, y)$ is true." From all this we can see that whereas primitive recursion corresponds to a **for** loop, μ-recursion corresponds to a **while** loop, with $g(\ldots)$ as the test condition.

EXAMPLE 3.5 *Partial square-root function.* Define the predicate $g(x, y)$ to hold if and only if $x = y^2$. Then the function f defined for all x by $f(x) = \mu y.g(x, y)$ computes the square root of x when x is a perfect square. When x is not a perfect square, however, the recursion is undefined, so f is a partial recursive function.

EXAMPLE 3.6 *Linear programming.* The standard *simplex algorithm* uses a **while** loop that executes a basic *pivot step* until a predicate expressing optimality holds. Hence the function that embodies the solution to a linear programming problem is μ-recursive. In point of fact, because a bound on the number of polytope corner points is explicitly definable from the problem instance, the same function can be computed via a simple **for** loop, so it is primitive recursive. However, the former method is usually much faster.

Part (a) of the next theorem expresses the fact that **while** loops, together with **if-then-else**, suffice to make a general-purpose programming language. Part (b) is the gist of the famous theorem, credited in various forms to various sources, that *at most one* **while** *loop is needed in any program.*

THEOREM 3.2 *(a) A (partial) function is μ-recursive if and only if it is a Turing-computable (partial) function.*

(b) Moreover, given any Turing machine T, we can find a primitive recursive function u and a primitive recursive predicate t such that for all x, $T(x) = u(\mu y.t(x, y))$.

In the standard proof of part (b), the predicate $t(x, y)$ is designed to hold if and only if y encodes the sequence of configurations of a halting computation of T on input x, and the function u picks off the output from the final configuration. To complete the proof of (a), all one needs to show is that given a Turing machine that computes g in Definition 3.2, one can build a Turing machine that computes f. This is done by following the unbounded-search procedure sketched above.

The corresponding theorem for formal languages also merits mention here. In Chapter 30 we defined the *characteristic function* of a language L to be the function f_L defined for all x by $f_L(x) = 1$ if $x \in L$, $f_L(x) = 0$ if $x \notin L$. This is simply the predicate corresponding to membership in L. The *partial characteristic function* still takes the value 1 when $x \in L$, but is undefined when $x \notin L$.

THEOREM 3.3 *(a) A language is recursive if and only if its characteristic function is μ-recursive.*

(b) A language is r.e. if and only if its partial characteristic function is μ-recursive.

Part (a) explains how the term "recursive" became applied to languages and predicates as a synonym for "decidable." It is important to recall that not all languages L accepted by Turing machines have computable characteristic functions (i.e., are decidable); unless we find a Turing machine accepting L that halts for all inputs, all we know is that the *partial* characteristic function of L is (partial-) computable. Before proceeding to *un*decidable languages, we take time to interpret these two theorems and others presented in the two previous chapters.

4 Equivalence of Computational Models and the Church-Turing Thesis

In Chapter 30 we introduced various machine models, the most important of which is the Turing machine. In Chapter 31 we introduced the grammar hierarchy of Chomsky, of which the most powerful was the Type-0 grammar. Here we have presented the purely mathematical model of μ-recursive functions. Although these models were defined over different domains for different purposes, they are all equivalent in a precise technical sense—they all define the same class of computable functions and decidable languages, and the same class of partial computable functions and partially decidable languages.

We can summarize all this by saying that Turing machines, type-0 grammars, and μ-recursive functions have the same problem-solving power.

This equivalence extends to vastly many other computational models, of which we mention a few:

(1) *Cellular Automata.* Cellular automata are intended to model the evolution of a colony of micro-organisms. Each cell is a deterministic finite automaton that receives its input in discrete time steps from neighboring cells, so that its current state is defined by its own previous state and the previous states of its neighbors. All the cells execute the same DFA. There are different schemes for specifying the representation of the input to a cellular automaton and its output. But under any reasonable scheme, the largest class of problems that can be solved on cellular automata coincides with the class of solvable problems on a Turing machine.

(2) *String-Rewriting Systems* A string-rewriting system is similar to a grammar. The main difference is that there are no non-terminals. Let the input alphabet be Σ. The production rules of a rewriting system T will be of the form $\alpha \rightarrow \beta$ where α and β are strings over Σ.

One can apply such a rule by replacing any occurrence of α in a string by β. T is defined as a finite set of rewrite rules, along with a finite set of initial strings. The language generated by T is defined as the set of strings that can be obtained from an initial string by applying the rewrite rules a finite number of times. The systems proposed before 1930 by Thue and Post fall roughly into this category. It turns out that the class of string-rewriting languages is the same as the r.e. languages (see [Book, 1993]).

(3) *Tree-Rewriting Systems.* These are similar to string-rewriting systems except that the local edits are done on subtrees of a tree, and rules may have more than one argument. The subtrees typically represent *terms* in algebraic or logical expressions that are being operated on. Under reasonable schemes for encoding numbers or strings by trees, all known tree-rewriting systems generate r.e. languages or compute partial recursive functions. Church's λ-*calculus* and most formal systems of logic fall into this category.

(4) *Extensions of Turing's Model.* As mentioned in Chapter 30, one can also create numerous modifications to the basic Turing machine model, such as having multi-dimensional tapes or binary trees with MOVE UP, MOVE DOWN LEFT, and MOVE DOWN RIGHT instructions (the latter are tantamount to having random-access to stored values), allowing nondeterminism or alternation, making computation probabilistic (see Chapter 35, Section 2), and so on. All of these machines compute the same functions as the simple one-tape Turing machine.

(5) *Random-Access Machines and High-Level Programming Languages.* These can be mentioned in tandem because a RAM, as described in Chapter 30, is just an idealization of assembly or machine language. Every high-level language yet devised can be compiled into some machine language. Even the standard *Java Virtual Machine* is little more than a RAM, with some added handling of class objects via pointers that is not unlike the workings of a pointer machine, and some hooks to enable the host system to control physical devices and network communications. Without excessive effort one can extend the construction of a universal Turing machine in Section 2 to handle the case where P is a RAM program rather than a GOTO program. The registers of P can be simulated on the tape by adding one more tape symbol # and using strings of the form $\#i\#j\#$, where i is the register's number and j is its contents. As stated in Chapter 30, Section 4, this simulation is even fairly efficient. Hence all

these high-level languages have the same problem-solving power as the lowly one-tape Turing machine.

The convergence of so many disparate formal models on the same class of languages or functions is the main evidence for the assertion that they all exactly capture the informal notion of what is mechanically or humanly computable. This assertion is called the **Church-Turing thesis**. In one form, it asserts that every problem that is humanly solvable is solvable by a Turing machine. Put more precisely, any cognitive process that a human being could or will ever use to distinguish certain numbers or strings as "good" defines an r.e. language—and if it also would determine that any other given number or string is "bad," it defines a recursive language. An extension of the thesis claims that no one will ever design a physical device to compute functions that are not μ-recursive. The Church-Turing thesis is not a mathematical conjecture and is not subject to mathematical proof; it is not even clear whether the extension is resolvable scientifically.

5 Undecidability

The Church-Turing thesis implies that if a language is undecidable in the formal sense defined above, then the problem it represents is really, humanly, physically undecidable. The existence of languages that are not even partially decidable can be established by a counting argument: Turing machines can be counted $1, 2, 3, \ldots$, but Cantor proved that the totality of all sets of integers cannot be so counted. Hence there are sets left over that are not accepted, let alone decided, by any program. This argument, however, does not apply to languages or problems *that one can state*, since these are also countable. The remarkable fact is that many easily-stated problems of high practical relevance are undecidable. This section shows that the five remaining problems on our list in Section 2.1, namely 2–5 and 7, are all unsolvable.

5.1 Diagonalization and Self-reference

Undecidability is inextricably tied to the concept of *self-reference*, and so we begin by looking at this perplexing and sometimes paradoxical concept. The simplest examples of self-referential paradox are statements such as "This statement is false" and "Right now I am lying." If the former statement is true, then by what it says, it is false; and if false, it is true. . . The idea and effects of self-reference go back to antiquity; a version of the latter "liar" paradox ascribed to the Cretan poet Epimenides even found its way into the New Testament, Titus 1:12–13. For a more-colorful example, picture a barber of Seville hanging out an advertisement reading, "I shave those who do not shave themselves." When the statement is applied to the barber himself, we need to ask: Does he shave himself? If *yes*, then he is one of those who do shave themselves, which are not the people his statement says he shaves. The contrary answer *no* is equally untenable. Hence the statement can be neither true nor false (it may be good ad copy), and this is the essence of the paradox. Such paradoxes have made entry into modern mathematics in various forms. We will present some more examples in the next few paragraphs. Many variations on the theme of self-reference can be found in the books of the logician and puzzlist Raymond Smullyan, including [Smullyan, 1978] and [Smullyan, 1992].

Berry's paradox concerns English descriptions of natural numbers. For example, the number 24 can be described by many different phrases: "twenty-four," "six times four," "four factorial," etc. We are interested in the *shortest* of such descriptions, namely one(s) having the fewest letters. Here, "two dozen" beats all of the above. Clearly there are (infinitely) many positive integers whose shortest descriptions require one hundred letters or more. (A simple counting argument can be used to show this. The set of positive integers is infinite, but the set of positive integers with English descriptions of fewer than one hundred letters is finite.) Let D denote the set of positive integers that do not have English descriptions of fewer than one hundred letters. Thus D is not empty. It is a well-known fact in set theory that any nonempty subset of positive integers has a smallest integer. Let x be the smallest integer in D. Does x have an English description of fewer than one hundred letters? By the definition of the set D and x, the answer is *yes*: such a description of x is, "the smallest positive integer that cannot be described in English in fewer than one hundred letters." This is an absurdity because the quoted part of the last sentence is clearly a description of x, and it contains fewer than one hundred letters.

Russell's paradox similarly turns on issues in defining sets. In formal mathematics, we can perfectly easily describe "the set of all sets that do not include themselves as elements" by the definition $S = \{x | x \notin x\}$. The question "Is $S \in S$?" leads to a real conundrum. This also resembles the barber paradox, with "$\notin$" read as "does not shave." This paradox forced the realization that the formal notion of a *set*, and importantly the formal rules that apply to sets, do not and cannot apply to everything that we informally regard as being a "set."

Our last example is a charming paradox named for the mathematician William Zwicker. Consider the collection of all two-person games that are *normal* in the sense that every play of the game must end after a finite number of moves. Tic-tac-toe is normal since it always ends within nine moves, while chess is normal because the official "fifty move rule" prevents games from going on forever. Now here is *hypergame*. In the first move of hypergame, the first player calls out a normal game—and then the two players go on to play that game, with the second player making the first move. Now we need to ask, "Is hypergame normal?" If *yes*, then it is legal for the first player to call out "hypergame!"—since it is a normal game. By the rules, the second player must then play the first move of hypergame—and this move can be calling out "hypergame!" Thus the players can keep saying "hypergame" without end, but this contradicts the definition of a normal game. On the other hand, suppose hypergame is not normal. Then in the first move, player 1 cannot call out hypergame and must call a normal game instead—so that the infinite move sequence given above is not possible and hypergame is normal after all!

Let us try to implement Zwicker's paradox. To play hypergame, we need a way of formalizing and encoding the rules of a game as a string x, and we need a decision procedure `isNormal(x)` to tell if the game is normal. Then the rules of hypergame are easily formalized: pick a string x, verify `isNormal(`x`)`, and play game x. Let h be the string encoding of these rules. Now we get a real contradiction when `isNormal(`h`)` is run. We must conclude that either (i) our formalization of games is inadequate or inconsistent, or (ii) a decision procedure `isNormal` simply cannot exist. Now (i) is the way out for Russell's paradox with "sets" in place of "games." For *computation*, however, we know that our formalization is adequate and consistent—and hence we will be faced with conclusion (ii), namely that our corresponding computational problems are unsolvable.

Before showing how the above paradoxes can be modified and ingrained into our problems, we need to review the 0-1 encoding of GOTO programs from Section 2.2, including the conventions that ACCEPT has the same code 100 as STOP for programs that accept languages, and that such an ACCEPT statement be last and unique. We may assign the code 101 to REJECT, which may appear anywhere. If a binary string x encodes a program P, it is easy to decode x into P, and we may identify x with P. If x does not encode a legal GOTO program, this fact is easy to detect.

Then we may choose to treat x as an alternate code for the trivial GOTO program that consists of a single REJECT statement.

Now we can define the so-called "diagonal language" L_d as follows:

$$L_d \;=\; \{x \mid x \text{ is a GOTO program that does not accept the string } x\} \tag{8}$$

This language consists of all programs in GOTO language that do not halt in the ACCEPT statement when given their own encoding as input—they may either REJECT or not halt at all on that input. For example, consider $x = 01111101101100$, which encodes a program that accepts any string beginning with 1 and rejects any string beginning with 0. Then $x \in L_d$ since the program does not accept 01111101101100. Note the self-reference in (8). Although the definition of L_d seems artificial, its importance will become clear in the next section when we use it to show the undecidability of other problems. First we prove that L_d is not even accepted by any Turing machine, let alone decided by one.

THEOREM 5.1 *L_d is not recursively enumerable.*

Proof. Suppose for the sake of contradiction that L_d is r.e. Then there is a GOTO program that accepts L_d—call it P. Now what does P do on input $x = code(P)$? If P accepts x, then x is not in L_d, but this contradicts $L(P) = L_d$. But if P does not accept x, then x *is* in L_d, and this also contradicts $L(P) = L_d$. Hence a program P such that $L(P) = L_d$ cannot exist, and so L_d is not r.e.

The definition of L_d is motivated by Russell's paradox, reading "$\notin$" as "does not accept." Whereas in Russell's paradox we had to conclude that S is not a *set*, here we conclude that L_d is not a *Turing-acceptable* set.

We can similarly carry over Zwicker's paradox by treating a given string x as formally defining "Game-x" as follows: The first player decodes x into a GOTO program P, and then tries to choose some string x' in the language $L(P)$. If $L(P)$ is empty, in particular if x decodes to the trivial program "1. REJECT" as stipulated above, then the game ends then and there. But if the first player finds such an x', then the second player must play the same way with x'. Then we can say that x is *normal* if every play of Game-x must terminate (by reaching a GOTO program that accepts the empty language) in a finite number of steps. Finally define L_Z to be the set of normal strings. By applying the reasoning from Zwicker's paradox, one can imitate the above proof to show that L_Z is not recursively enumerable.

5.2 Reductions and More Undecidable Problems

Recall from Chapter 30 (section 2.3) the notion of Turing reducibility. Basically, a language L_1 is Turing reducible to L_2 if there is a halting Turing machine for language L_1 using an oracle for language L_2. If L_1 is reducible to L_2 and L_2 is decidable, then so is L_1. Basically, one can replace queries to oracles by executing a halting computation for L_2. The contrapositive of this statement can be used to show undecidability. If L_1 is undecidable, then so is L_2. We will first express Problem (2) as a language:

$L_U = \{code(P)111x |\ P$ accepts the string $x\}$.

Thus L_U takes as input a program in GOTO, and a binary string x, and accepts the encoded pair (P, x) if and only if P accepts x. (Note 111 is used as a separator between P and x.) The universal program presented in Section 2.2 accepts the language L_U hence it is recursively enumerable. We will show that L_U is not recursive. First, we will show a simple fact about recursive languages.

THEOREM 5.2 *Recursive languages are closed under complement.*

Proof. Let P be a GOTO program for language L. The program P' obtained by exchanging ACCEPT and REJECT instructions is easily seen to accept the language $\bar{L}$. This standard trick works to complement the computations of most of the deterministic devices (such as DFA). •

Now we show that L_U is not recursive.

THEOREM 5.3 *L_U is not recursive.*

Proof. Consider the language $L'_U = \{x |\ x$ when interpreted as a GOTO program accepts its own encoding$\}$. Obviously, $L'_U = \bar{L}_D$. Since L_D is not recursively enumerable, it is not recursive. (Recall that the set of recursive languages is a subset of recursively enumerable languages.) By the above theorem, L'_U is not recursive. Finally, note that L_U can be reduced to L'_U as follows: Given an algorithm for L_U, we can construct an algorithm for L'_U as follows: Let P be an algorithm for L_U. To construct an algorithm for L'_U simply note the connection between the two problems. An input string x belongs to L'_U if and only if $x111x$ belongs to L_U. Thus, a simple copy program (similar to one presented in Section 2) can be first used to convert the input x into $x111x$. Move the scanning head back to the leftmost character of the first copy of x. Now simply run the program P. Note that the program P' described above is being constructed using only P, not x. This reduction shows that L_U is not recursive.

Next we consider problem (6) in our list. Earlier we showed that a special case of this problem (when the input is restricted to type-1 grammar) is totally solvable. It is not hard to see that the general problem is partially solvable. (To see this, suppose there is a derivation for a string x starting from S, the start symbol of the grammar. Suppose the length of one such derivation is k. A program can try all possible derivations of length 1, 2, etc. until it succeeds. Such a program will always halt on strings x generated by the grammar G. Thus the language

$\{< G > \#x| < G >$ is a type-0 grammar and x can be generated by G $\}$

is recursively enumerable. A standard result from formal language theory [Hopcroft and Ullman, 1979] is that for every Turing machine M, there is a type-0 grammar such that $L(M) = T(G)$. This conversion is the reduction that shows that the language stated above is not recursive.

By a more elaborate reduction (from L_D), it can be shown that tiling problem (4) in our list is also not partially decidable. We will not do it here and refer the interested reader to [Harel, 1992]. But we would like to point out how the undecidability result can be used to infer a result about *aperiodic tilings*. This deduction is interesting because the result appears to have some deep implications and is hard to deduce directly. We need the following definition before we can state the result. A different way to pose the tiling problem is whether a given set of tiles can tile *an entire plane* in such a way that all the adjacent tiles have the same color on the meeting quarter. (Note that this question is different from the way we originally posed it: Can a given set of tiles tile any *finite* rectangular region? Interstingly, the two problems are identical in the sense that the answer to one version is “yes” if and only if it is “yes” for the other version.) Call a tiling of the plane it periodic if one can identify a $k \times k$ square such that the the entire tiling is made by repeating this $k \times k$ square tile. Otherwise, call it *aperiodic*. Consider the question: Is there a (finite) set of unit tiles which can tile the plane, but only aperiodically? The answer is “yes” and it can be shown by from the total undecidability of tiling problem. Suppose the answer is "no". Then, for any given set of tiles, the entire plane can be tiled if and only if the plane can be tiled periodically. But a periodic tiling can be found, if one exists, by trying to tile a $k \times k$ region for successively increasing values of k. This process will eventually succeed (in a finite number of steps) if the tiling exists. This will make the tiling problem partially decidable, which contradicts the total undecidability of the problem. This means that the assumption that the entire plane can be tiled if and only if some $k \times k$ region can be tiled is wrong. Thus there exist a (finite) set of tiles that can tile the entire plane, but only aperiodically.

We conclude with a brief remark about problem (7) in our list. After many years of effort by several mathematicians and computer scientists (including Davis and Robinson), Matiyasevich found an effective way to transform a given Turing machine T into a set of equations in variables $x, y_1, \ldots, y_m$ such that for any x, T on input x halts if and only if the other m variables can be set to solve the equations. This reduction shows that Hilbert's tenth problem is undecidable. Details behind this reduction can be found in [Floyd and Beigel, 1994].

6 Defining Terms

Decision Problem: A computational problem with yes/no answer. Equivalently a function whose range consists of two values $\{0, 1\}$.

Decidable Problem: A decision problem which can be solved by a GOTO program that halts on all inputs in a finite number of steps. For emphasis, the equivalent term *totally decidable* problem is used. The associated language is called recursive.

Partially Decidable Problem: A decision problem that can be solved by a GOTO program which halts (and outputs ACCEPT) on all yes instances. The program may or may not halt on no instances. Equivalently, the collection of yes instance strings forms a type-0 language. (See Chapter 31.)

Recursively Enumerable Language: Same as partially decidable language. This term was not used in this chapter.

μ-Recursive Function: A function that is a basic function (Zero, Successor or Projection), or one that can be obtained from other μ-recursive functions using composition and μ-recursion.

Recursive Language: A language that can be accepted by a GOTO program that halts on all inputs. The associated problem is called decidable.

Solvable Problem: A computational problem which can be solved by a halting GOTO program. The problem may have a non-binary output.

Totally Undecidable Problem: A problem which cannot be solved by a GOTO program. Equivalently, the set of yes instance strings is not a type-0 language.

Undecidable Problem: A decision problem that is not (totally) decidable. It could be partially decidable or totally undecidable.

Universal Turing Machine: A Turing machine that can simulate any other Turing machine.

Unsolvable Problem: A computational problem which is not solvable. The associated function will be called uncomputable function.

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Further Information

The fundamentals of computability can be found in many books including the classical books [Rogers, 1967, Davis, 1958, Minsky, 1967]. More recent books on automata and formal languages have also devoted at least a few chapters on computability [Floyd and Beigel, 1994, Gurari, 1989, Harel, 1992, Harrison, 1978, Hopcroft and Ullman, 1979, Sipser, 1996, Wood, 1987]. Early work on computability was motivated by a quest to address profound questions about the basis of logical reasoning, mathematical proofs and automatic computation. Various formalisms discussed in this chapter were proposed at around the same time, and soon thereafter, their equivalence was tested. Thus, in a short time, Church thesis took deep roots. Subsequent work has focussed on resolution of specific problems (as to decidable or not). Two such famous problems were Hilbert's tenth problem and word problem for groups. Both problems were eventually shown to be unsolvable. Another direction of research has been to make finer distinction among unsolvable problems by introducing degrees of unsolvability. Recursive function theory and lambda calculus also led to the development of functional programming languages (such as Lisp, Scheme, Haskell, ML etc.) Computability theory is also closely related to logic, formal deductive systems and structural complexity. Logic and deductive systems are of interest to philosophers and researchers in artificial intelligence, as well as computation theorists. Although there are no journals devoted exclusively to computability, many theory journals (e.g. listed at the end of the last two chapters) publish papers on this topic. In addition, Annals of Pure and Applied Logic publishes papers on logic and computability.