

COMPLEXITY THEORY

Lecture 4: Undecidability and Recursion

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Rice's Theorem

We can make this formal as follows:

Definition 4.2: Let \mathcal{P} be a set of languages. A language L has the property \mathcal{P} if $L \in \mathcal{P}$. Property \mathcal{P} is a **non-trivial** property of recognisable languages if there are TM-recognisable languages that have it and others that do not have it.

Theorem 4.1 (Rice's Theorem): If \mathcal{P} is a non-trivial property of recognisable languages, then the following problem is undecidable:

$$\mathcal{P}\text{-ness} = \{\langle M \rangle \mid L(M) \in \mathcal{P}\}$$

Undecidability so far

We have seen several undecidable problems for TMs:

- The **Halting Problem**: recognise TM-word pairs where the TM halts
- The **Non-Halting Problem**: recognise TM-word pairs where the TM does not halt
- The **ε -Halting Problem**: recognise TMs that halt on the empty input

Many further TM-related problems are undecidable . . .

. . .but we can use a shortcut to proving many of them:

Theorem 4.1 (Rice's Theorem, informal): Any interesting property related to the language recognised by a given TM is undecidable.

Proof of Rice's Theorem

Theorem 4.1 (Rice's Theorem): If \mathcal{P} is a non-trivial property of recognisable languages, then the following problem is undecidable:

$$\mathcal{P}\text{-ness} = \{\langle M \rangle \mid L(M) \in \mathcal{P}\}$$

Proof: We reduce ε -Halting to \mathcal{P} -ness.

- Assume w.l.o.g. that $\emptyset \notin \mathcal{P}$ (otherwise do the proof for $\overline{\mathcal{P}}$)
- Let M_L be some TM that recognises a language $L \in \mathcal{P}$
- Given any TM M , compute a TM M^* that behaves as follows:
On input $w \in \Sigma^*$: (1) Simulate M on input ε
(2) If M halts, simulate M_L on w
- Then $L(M^*) = L \in \mathcal{P}$ if M halts on ε , and
 $L(M^*) = \emptyset \notin \mathcal{P}$ if M does not halt on ε

For the required Turing reduction, we construct a TM that:

(Step 1) checks if the input is a TM encoding $\langle M \rangle$ and rejects otherwise,

(Step 2) returns the result of the check $\langle M^* \rangle \in \mathcal{P}$. This would decide ε -Halting. \square

Using Rice's Theorem

Here are some simple results that Rice gives us:

Corollary 4.3: Given an arbitrary TM \mathcal{M} , it is undecidable whether the language recognised by \mathcal{M} has any of the following properties:

- emptiness
- finiteness
- decidability
- regularity
- context-freedom
- contains any given word w (word problem for TMs)

Attention: There are of course many non-trivial properties of TMs that can be decided, and which do not relate to their language:

Example 4.4: It is decidable if a TM has at least three states.

Semi-decidable + Co-semi-decidable = Decidable

An easy but important observation:

Theorem 4.7: If L is semi-decidable and co-semi-decidable, then L is decidable.

Proof: On input w , simulate, in parallel, a recogniser for L and a recogniser for \bar{L} . At least one of them eventually must halt, so we can decide if $w \in L$. \square

We thus obtain an example of a problem that is not Turing-recognisable.

Corollary 4.8: The Non-Halting Problem is not Turing-recognisable.

Semi-decidability and Co-semi-decidability

We can distinguish the following two cases:

- (1) L is Turing-recognisable: L is semi-decidable
- (2) \bar{L} is Turing-recognisable: L is co-semi-decidable

We have seen examples for both:

Theorem 4.5: The Halting Problem is semi-decidable.

Proof: Use the universal TM to simulate an input TM, and accept if it halts. \square

Corollary 4.6: The Non-Halting Problem is co-semi-decidable.

Turing reductions and semi-decidability

Observation:

- If Q is decidable and $P \leq_T Q$, then P is decidable (Theorem 3.17)
- But: if Q is semi-decidable and $P \leq_T Q$, then P may or may not be semi-decidable

Reason: An oracle for Halting is as good as an oracle for Non-Halting, since we are free to complement the answer in an oracle machine.

This is a general insight: complementing oracles has no effect

To preserve (co-)semi-decidability, one needs a more restricted form of reduction:

Definition 4.9: A language P is *many-one reducible* to a language Q , written $P \leq_m Q$ if there exists a total computable function $f : \Sigma^* \rightarrow \Sigma^*$ such that, for all $w \in \Sigma^*$:

$$w \in P \quad \text{if and only if} \quad f(w) \in Q.$$

This is sometimes called a *mapping-reduction* or an *m-reduction*.

Properties of Many-One-Reductions

Many-one reductions are special kinds of Turing reductions:

Theorem 4.10: If $P \leq_m Q$ then $P \leq_T Q$.

Proof: Given a TM that recognises Q , we obtain a TM that recognises P as follows:

- On input w , compute $f(w)$
- Simulate the TM for Q and return the result (if any) □

An easy consequence of Theorem 3.17 therefore is:

Corollary 4.11: If $P \leq_m Q$ and Q is decidable, then P is decidable.

However, now we also have the following:

Theorem 4.12: If $P \leq_m Q$ and Q is semi-decidable, then P is semi-decidable.

Proof: Use the same TM as in the proof of Theorem 4.10. □

Equivalence is Hard

We can show a somewhat stronger result:

Theorem 4.15: Equivalence of Turing machines is neither semi-decidable nor co-semi-decidable.

Proof: We have already shown ε -Halting \leq_m Equivalence. Since we know that ε -Halting is not co-semi-decidable (similar to Halting), we conclude that Equivalence is neither.

However, we can also show that $\overline{\varepsilon}$ -Halting \leq_m Equivalence.

- Note that the TM \mathcal{M}^* defined on the previous slide either accepts all inputs (if \mathcal{M} halts on ε) or none (if it doesn't)
- Equivalence to \mathcal{M}_ε corresponds to ε -Halting
- On the other hand, equivalence to a TM \mathcal{M}_\emptyset , which rejects all inputs, corresponds to ε -non-Halting

We can therefore use the reduction f :

$$f(w) = \begin{cases} \langle \mathcal{M}^*, \mathcal{M}_\emptyset \rangle & \text{if } w = \langle \mathcal{M} \rangle \\ \varepsilon & \text{(an invalid input) if } w \text{ is no encoded TM} \end{cases} \quad \square$$

Example: Many-one Reduction

Some of our previous Turing-reductions can easily be described as many-one, e.g., Halting can be many-one reduced to ε -Halting. Here is another example:

Definition 4.13: Two TMs \mathcal{M} and \mathcal{N} are **equivalent** if $L(\mathcal{M}) = L(\mathcal{N})$.

Theorem 4.14: Equivalence of Turing machines is undecidable.

(Note that we could also get this from Rice's Theorem, but we want to try out our new machinery.)

Proof: We define f such that $w \in \varepsilon$ -Halting iff $f(w) \in$ Equivalence.

Let \mathcal{M}_ε be a TM that accepts all inputs.

For a TM \mathcal{M} , we define the following TM \mathcal{M}^* :

- Simulate \mathcal{M} on the empty input.
- If \mathcal{M} halts, accept.

Then \mathcal{M}^* is equivalent to \mathcal{M}_ε iff \mathcal{M} halts on the empty input. We define f :

$$f(w) = \begin{cases} \langle \mathcal{M}^*, \mathcal{M}_\varepsilon \rangle & \text{if } w = \langle \mathcal{M} \rangle \\ \varepsilon & \text{(an invalid input) if } w \text{ is no encoded TM} \end{cases} \quad \square$$

Recursion

A Paradox

A Paradox in the Study of Life:

- (1) Living things are machines.
- (2) Living things can reproduce.
- (3) Machines cannot reproduce.

Rationale:

- (1) Viewpoint of modern biology.
- (2) Clear.
- (3) If a machine A produces a machine B , then A must be more complex than B . For example, a car-producing factory is **more complex** than the cars it produces, as it contains the design of the cars and, **in addition**, the design of all manufacturing robots, among others. Since no machine is more complex than itself, a machine cannot reproduce itself.

Resolving the Paradox

A Paradox in the Study of Life:

- (1) Living things are machines.
- (2) Living things can reproduce.
- (3) Machines cannot reproduce.

Question: How to resolve this paradox?

Answer: Assertion (3) is wrong.

In particular, the underlying argument of “more information” and “greater complexity” needed by the producing machine is flawed: there are TMs that reproduce themselves

Quines

Reproduction of TMs is closely related to the task of creating a program that prints its own source code:

Definition 4.16: A **quine** is a program that, when started without any input, will print out its own source code, and then stop.

Can Quines be created? How?

Example 4.17 (A quine in English): Print this sentence.

However, we cannot turn this into a program, since “this sentence” does often not correspond to available programming constructs.

Example 4.18 (Another quine in English): Print the following sentence twice, the second times in quotes. "Print the following sentence twice, the second times in quotes."

Some Real Quines

Example 4.19 (A classic C quine): `main()char *c="main()char *c=%c%s%c;printf(c,34,c,34);";printf(c,34,c,34);`

Example 4.20 (The shortest C quine, by Szymon Rusinkiewicz):

Example 4.21 (A Python quine by Frank Stajano):
`l='l=%s;print l%'l';print l%'l'`

Note: A variation are **ouroboros quines** that print out another program that prints out the original again. More steps are possible. See, e.g., <https://github.com/mame/quine-relay> for one with 100 steps.

Other variations exist (see Wikipedia).

Towards a TM Quine

We define a TM SELF that ignores its input and prints out a description of itself.
(A TM quine, where “source code” is interpreted as “encoding of the TM”)

The following small result is helpful:

Lemma 4.22: There is a computable function $q : \Sigma^* \rightarrow \Sigma^*$ such that, for each $w \in \Sigma^*$, the word $q(w)$ is a TM that prints w and halts.

Proof: For any word w , let \mathcal{P}_w be a TM that replaces the tape contents with the word w (clearly, this can easily be found for any w)

Now q is simply computed by a TM that, given w as input, constructs \mathcal{P}_w and then computes and outputs $\langle \mathcal{P}_w \rangle$. □

Summary: SELF TM

So how did we construct our TM quine now?

Step 1: We define some TM B that behaves as follows:

Given some input string $\langle M \rangle$:

- compute $q(\langle M \rangle)$
- concatenate the TMs given by $q(\langle M \rangle)$ and $\langle M \rangle$
(take a disjoint union of states where any halting state of $\langle M \rangle$ gets a transition to the starting state of $q(\langle M \rangle)$)
- output the encoding of the resulting machine

Step 2: We define SELF to be the TM constructed by B on input $\langle B \rangle$

Defining the TM SELF

Like other quines, SELF consists of two parts:

A Compute the “source code” $\langle B \rangle$ of a suitable program B

B Use $\langle B \rangle$ to print out:

- (1) source code $\langle A \rangle$ that computes $\langle B \rangle$ and (2) the source code $\langle B \rangle$ itself

We know how to implement part A: use the TM $\mathcal{P}_{\langle B \rangle}$

(however, to actually do this, we need to know B first)

B in turn can work as follows:

Given some input string $\langle M \rangle$:

- compute $q(\langle M \rangle)$
- concatenate the TMs given by $q(\langle M \rangle)$ and $\langle M \rangle$
(take a disjoint union of states where any halting state of $\langle M \rangle$ gets a transition to the starting state of $q(\langle M \rangle)$)
- output the encoding of the resulting machine

Then part B does not depend on A , so we can really define A as $\mathcal{P}_{\langle B \rangle}$

The Recursion Theorem

Going further, we can allow any TM to access its own description during the computation:

Theorem 4.23 (Recursion Theorem): Let $t : \Sigma^* \times \Sigma^* \rightarrow \Sigma^*$ be a function computed by some TM \mathcal{T} (assuming a suitable encoding of pairs of words over Σ^*). Then there is a TM \mathcal{R} that computes a function $r : \Sigma^* \rightarrow \Sigma^*$ such that

$$r(w) = t(\langle \mathcal{R} \rangle, w)$$

for every $w \in \Sigma^*$.

Intuition: To make a TM that can use its own description, we first devise a TM \mathcal{T} (to compute t) that receives the description of a machine as extra input. The theorem yields a TM \mathcal{R} that operates like \mathcal{T} does but with \mathcal{R} 's description filled in automatically.

The Recursion Theorem: Proof

Theorem 4.23 (Recursion Theorem): Let $t : \Sigma^* \times \Sigma^* \rightarrow \Sigma^*$ be a function computed by some TM \mathcal{T} (assuming a suitable encoding of pairs of words over Σ^*). Then there is a TM \mathcal{R} that computes a function $r : \Sigma^* \rightarrow \Sigma^*$ such that $r(w) = t(\langle \mathcal{R} \rangle, w)$ for every $w \in \Sigma^*$.

Proof: The proof is similar to the construction of SELF, using a TM with three parts A , B and T :

- A : print $\langle BT \rangle$ (like $\mathcal{P}_{\langle BT \rangle}$ but without deleting the input)
we use BT to denote the concatenation of the TM parts B and T in one TM
- B : on an input of form $w\langle M \rangle$, replace $\langle M \rangle$ by an encoding of the concatenation of $q'(\langle M \rangle)$ and $\langle M \rangle$
where $q'(v)$ is like q but returns a TM that adds v at the end of the tape
- T : run \mathcal{T} on an input of form $w\langle N \rangle$

We assume that our TM encoding can be written next to the input w without risk of confusion. Then \mathcal{R} is the TM obtained as the concatenation of A , B , and T

This is the TM whose encoding B would write on some input $w\langle BT \rangle$ □

Halting is Undecidable: Proof by Introspection

We can also use the Recursion Theorem for alternative proofs:

Theorem 3.11 The Halting Problem \mathbf{P}_{Halt} is undecidable.

Proof: By contradiction: Suppose there is a decider \mathcal{H} for the Halting Problem

We construct a TM \mathcal{M} that, on input w , acts as follows:

- (1) Obtain own description $\langle \mathcal{M} \rangle$
- (2) Simulate \mathcal{H} on input $\langle \mathcal{M} \rangle \# \langle w \rangle$, that is, check if \mathcal{M} halts on w
- (3) If yes, enter an infinite loop;
if no, halt and accept

Then \mathcal{M} halts on w if and only if it doesn't – contradiction. □

Using the Recursion Theorem

By the Recursion Theorem, we can now use instructions like “obtain own description $\langle \mathcal{M} \rangle$ ” in our informal descriptions of TMs.

Example 4.24: We can describe a TM quine in the style of our previous SELF as follows:

On any input:

- Obtain own description $\langle \mathcal{M} \rangle$
- Print $\langle \mathcal{M} \rangle$

We can construct such a TM by applying the Recursion Theorem to the TM \mathcal{T} described as follows:

On input $\langle w, \mathcal{M} \rangle$, print $\langle \mathcal{M} \rangle$

The Recursion Theorem turns this into a TM \mathcal{R} that is a quine.

Minimal TMs

Definition 4.25: A TM \mathcal{M} is called **minimal** if there is no TM equivalent to \mathcal{M} that has a shorter description. The problem of deciding if a TM is minimal is:

$$\mathbf{MIN}_{\text{TM}} = \{ \langle \mathcal{M} \rangle \mid \mathcal{M} \text{ is a minimal TM} \}$$

Theorem 4.26: \mathbf{MIN}_{TM} is not Turing-recognisable.

Proof: Assume there is some TM \mathcal{E} enumerating \mathbf{MIN}_{TM} .

We define a TM \mathcal{C} that processes an input w as follows:

- (1) Obtain own description $\langle \mathcal{C} \rangle$
- (2) Simulate \mathcal{E} until some TM \mathcal{D} is printed such that $\langle \mathcal{D} \rangle$ is longer than $\langle \mathcal{C} \rangle$
- (3) Simulate \mathcal{D} on w

Then \mathcal{C} is equivalent to \mathcal{D} , but it has a shorter description, contradicting the assumption that \mathcal{D} is minimal. □

Summary and Outlook

Most properties related to the computation of TMs are undecidable

Many-one reductions establish a closer relationship between two problems than Turing reductions

There are non-semi-decidable problems

Turing machines can work with their own description

What's next?

- No lectures next week
- Defining complexity classes
- Time complexity