Theorem 3.16 Every nondeterministic Turing machine has an equivalent deterministic Turing machine.

- A nondeterministic TM can be simulated with a 3-tape Turing machine, which goes systematically through the possible computations of the nondeterministic TM
- Tape 1 maintains the input
- Tape 2 simulates the tape of the nondeterministic TM
- Tape 3 keeps track of the situation with the possible computations

- The tree formed of the possible computation paths of the nondeterministic TM has to be examined using breadth-first algorithm, not in depth-first order
- Let $b$ be the largest number of possible successor states for any state in the Turing machine
- Each node in the tree can be indexed with a string over the alphabet $\{1, 2, \ldots, b\}$
- For example, $231$ is the node that is the first child of the third child of the second child of the root of the tree
- All strings do not correspond to legal computations (nodes of the tree) — those get rejected
- Going through the strings in the lexicographic order corresponds to examining all computation paths, and the search order in the tree is breadth-first search
Working idea:

- Copy the input from tape 1 to tape 2
- Tape 3 tells us which is (in the lexicographic order) the next computation alternative for the nondeterministic TM. We simulate that computation targeting the updates on the tape of the original TM into tape 2 of the simulating TM
- Observe that there is a finite number of transition possibilities (successor states)
- Systematic examination of the possible computations of the nondeterministic TM ends into the accepting final state only if the original TM has an accepting computation path
- If no accepting computation exists, the simulating TM never halts

UNRESTRICTED GRAMMARS

- Context-free grammar allows to substitute only variables with strings
- In an unrestricted grammar (or a rewriting system) one may substitute any non-empty string (containing variables and terminals) with another one (also with the empty string $\epsilon$)

An unrestricted grammar is a 4-tuple $G = (V, \Sigma, R, S)$, where

- $V$ is the set of variables,
- $\Sigma$ is the set of terminals,
- $\Gamma = V \cup \Sigma$ is the alphabet of $G$,
- $R \subseteq \Gamma^* \times \Gamma^*$ is the set of rules, and
- $S \in V$ is the start variable

$(w, w') \in R$ is denoted as $w \rightarrow w'$
Let

- $G = (V, \Sigma, R, S)$,
- strings $v \in \Gamma^*$ and $u, w, x \in \Gamma^*$ as well as
  - $v \rightarrow x$ a rule in $R$
- $uvw$ yields string $uxw$ in grammar $G$,
  
- $uvw \Rightarrow_G uxw$
- String $v$ derives string $w$ in grammar $G$,
  
- $v \Rightarrow_G w$,

  if there exists a sequence $v_1, v_2, \ldots, v_k \in \Gamma^*$ (k ≥ 0) s.t.
  
- $v \Rightarrow_G v_1 \Rightarrow_G v_2 \Rightarrow_G \ldots \Rightarrow_G v_k \Rightarrow_G w$
- $k=0$: $v \Rightarrow_G v$ for any $v \in \Gamma^*$

$u \in \Gamma^*$ is a sentential form of $G$ if $S \Rightarrow_G u$

- A sentential form consisting only of terminal symbols $w \in \Sigma^*$ is a sentence of $G$
- The language of the grammar $G$ consists of sentences
  
- $L(G) = \{ w \in \Sigma^* | S \Rightarrow_G w \}$

The language $\{ a^k b^k c^k | k \geq 0 \}$ is not a context-free one; it can be generated with an unrestricted grammar, which

1. Generates the variable sequence $L(ABC)^k$ (or $\varepsilon$)
2. Orders the variables lexicographically $\Rightarrow LA^k B^k C^k$
3. Replaces the variables with terminals
For example, we can generate the sentence \textit{aabbcc} as follows

\begin{align*}
S & \Rightarrow LT \Rightarrow LABCT \Rightarrow LABCABC \\
& \Rightarrow LABACBC \Rightarrow LAABCBC \\
& \Rightarrow LAABBCC \Rightarrow aABBCC \\
& \Rightarrow aaBBCC \Rightarrow aabBCC \\
& \Rightarrow aabBC \Rightarrow aabbcC \\
& \Rightarrow aabbc \\
& \Rightarrow aabbcc
\end{align*}
**Theorem** A formal language $L$ generated by an unrestricted grammar can be recognized with a Turing machine.

**Proof.** Let $G = (V, \Sigma, R, S)$ be the unrestricted grammar generating language $L$. We devise a two-tape nondeterministic Turing machine $M_G$ for recognizing $L$.

$M_G$ maintains the input string on tape 1. On tape 2 there is some sentential form of $G$ which we try to rewrite as the input string.

At the beginning tape 2 contains the start variable $S$.

The computation of $M_G$ repeats the following stages:

1. The tape head of tape 2 is (non-deterministically) moved to some location on the tape;
2. we choose (non-deterministically) some rule of $G$ and try to apply it to the chosen location of the tape;
3. if the symbols on the tape match the symbols on the left-hand side of the rule, $M_G$ replaces them on tape 2 with the symbols on the right-hand side of the rule;
4. we compare the strings in tapes 1 and 2 with each other;
   a) if they are equal, the Turing machine enters the accepting final state and halts,
   b) otherwise, we go back to step 1.
**Theorem.** If a formal language $L$ can be recognized with a Turing machine, then it can be generated with an unrestricted grammar.

**Proof.** Let $M = (Q, \Sigma, \Gamma, \delta, q_0, \text{accept}, \text{reject})$ be a standard Turing machine recognizing language $L$.

Let us compose an unrestricted grammar $G_M$ that generates $L$.

As variables of the grammar we take symbols representing all states $q \in Q$ of $M$. The configuration of the TM $u q a v$ is represented as string $[uqav]$.

By the transition function of $M$ we give $G_M$ rules so that $[uqav] \Rightarrow_G [u'q'a'v'] \Leftrightarrow u q a v \Rightarrow_M u'q'a'v'$.

Then $x \in L(M) \Leftrightarrow [q_0x] \Rightarrow_G [\text{accept}v], \; u, v \in \Sigma^*$.

There are three types of rules in $G_M$:

1. Those that generate any string $x[q_0x], x \in \Sigma^*$ and $[,\;], q_0 \in \Sigma$ from the start variable:

   - $S \rightarrow T[q_0]$
   - $T \rightarrow \varepsilon$
   - $T \rightarrow aTA_a$
   - $A_a[q_0] \rightarrow [q_0A_a]$
   - $A_a b \rightarrow bA_a$
   - $A_a \rightarrow a$
2. Those that simulate the transition function of the Turing machine:

\[
\begin{align*}
\delta(q, a) &= (q', b, R) & qa &\rightarrow bq' \\
\delta(q, a) &= (q', b, L) & cqa &\rightarrow q'cb \\
\delta(q, \square) &= (q', b, R) & q &\rightarrow bq' \\
\delta(q, \square) &= (q', b, L) & cq &\rightarrow q'cb \\
\delta(q, \square) &= (q', \square, L) & cq &\rightarrow q'c
\end{align*}
\]

3. Those that replace a string of the form \([\text{accept}\ v]\) to an empty string

\[
\begin{align*}
\text{accept} &\rightarrow E_L E_R \\
\alpha E_L &\rightarrow E_L \\
[E_L &\rightarrow \varepsilon \\
E_R a &\rightarrow E_R \\
E_R &\rightarrow \varepsilon
\end{align*}
\]

Now a string \(x\) in \(L(M)\) can be generated as follows

\[
S \Rightarrow_1 x[q_0 x] \Rightarrow_2 x[\text{accept}\ v] \Rightarrow_3 x
\]
3.3 The Definition of Algorithm

- The formulations of computation by Alonzo Church and Alan Turing were given in response to Hilbert’s tenth problem which he posed in 1900 in his list of 23 challenges for the new century
- What Hilbert essentially asked for was an algorithm for determining whether a polynomial has an integral root
- Today we know that this problem is algorithmically unsolvable
- It is possible to give algorithms without them being exactly defined, but it is not possible to show that such cannot exist without a proper definition
- It was not until 1970 that Matijasevič showed that no algorithm exists for testing whether a polynomial has integral roots

Expressed as a formal language Hilbert’s tenth problem is

\[ D = \{ p \mid p \text{ is a polynomial with an integral root} \} \]

- Concentrating on single variable polynomials we can see how the language \( D \) could be recognized
- In order to find the correct value of the only variable, we go through its possible integral values 0, 1, -1, 2, -2, 3, -3, ...
- If the polynomial attains value 0 with any examined value of the variable, then we accept the input
- A similar approach is possible when there are multiple variables in the polynomial
For a single variable polynomial the roots must lie within
\[ \pm k \left( \frac{c_{\text{max}}}{c_1} \right) , \]
where \( k \) is the number of terms in the polynomial,
\( c_{\text{max}} \) is the coefficient with largest absolute value, and
\( c_1 \) is the coefficient of the highest order term.

If a root is not found within these bounds, the machine rejects.
Matijasevič’s theorem shows that calculating such bounds for
multivariable polynomials is impossible.

The language \( D \) can, thus, be recognized with a Turing machine,
but cannot be decided with a Turing machine (may never halt).

**Computability Theory**

- We will examine the *algorithmic solvability* of problems
  - I.e. solvability using Turing machines

- We make a distinction between cases in which formal
  languages can be recognized with a Turing machine and
  those in which the Turing machine is required to halt with
  each input.

- It turns out that there are many natural and interesting
  problems that cannot be solved using a Turing machine.

- Hence, by Church-Turing thesis these problems are
  unsolvable by a computer!
Definition 3.5 Call a language Turing-recognizable (or recursively enumerable, RÉ-language) if some Turing machine recognizes it.

Definition 3.6 Call a language Turing-decidable (or decidable, or recursive) if some TM decides it (halts on every input, is total).

- The decision problem corresponding to language $A$ is decidable if $A$ is Turing-decidable.
- A problem that is not decidable is undecidable
- The decision problem corresponding to language $A$ is semidecidable if $A$ is Turing-recognizable
- Observe: an undecidable problem can be semidecidable.

Basic Properties of Turing-recognizable Languages

Theorem A Let $A, B \subseteq \Sigma^*$ be Turing-decidable languages. Then also languages
1. $\bar{A} = \Sigma^* \setminus A$,
2. $A \cup B$, and
3. $A \cap B$
are Turing-decidable.

Proof.

1. Let $M_A$ be the total TM recognizing language $A$. By exchanging the accepting and rejecting final state of $M_A$ with each other, we get a total Turing machine deciding the language $\bar{A}$.
2. Let $M_A$ and $M_B$, respectively, be the total Turing machines deciding $A$ and $B$.
   - Let us combine them so that we first check whether $M_A$ accepts the input.
   - If it does, so does the combined TM.
   - On the other hand, if $M_A$ rejects the input, we pass it on to TM $M_B$ check.
   - In this case $M_B$ decides whether the input will be accepted. $M_A$ must pass the original input to $M_B$.
   - It is clear that the combined TM is total and accepts the language

$$A \cup B = \{ x \in \Sigma^* | x \in A \lor x \in B \}$$

3. The Turing-decidability of $(A \cap B)$ follows from the previous results because

$$A \cap B = \overline{A} \cup \overline{B}$$
Theorem B Let $A, B \subseteq \Sigma^*$ be Turing-recognizable languages. Then also languages $A \cup B$ and $A \cap B$ are Turing-recognizable.

Proof. Exercises.

- As a consequence of Theorem 4.22 we get a hold of languages that are not Turing-recognizable ($\overline{A}$ is the complement of $A$):

Theorem C Let $A \subseteq \Sigma^*$ be a Turing-recognizable language that is not Turing-decidable. Then $\overline{A}$ is not Turing-recognizable.
Theorem 4.22  

Language $A \subseteq \Sigma^*$ is Turing-decidable $\iff$ $A$ and $\overline{A}$ are Turing-recognizable.

Proof.

\(\Rightarrow\) If $A$ is Turing-decidable, then it is also Turing-recognizable. By Theorem A(1) the same holds also for $\overline{A}$.

\(\Leftarrow\) Let $M_A$ and $M_{\overline{A}}$ be the TMs recognizing $A$ and $\overline{A}$.

For all $x \in \Sigma^*$ it holds that either $x \in A$ or $x \notin A$. In other words, either $M_A$ or $M_{\overline{A}}$ halts on input $x$.

Combining machines $M_A$ and $M_{\overline{A}}$ to run parallel gives a total Turing machine for recognizing $A$. $\Box$
Encoding Turing Machines

Standard Turing machine

\[ M = (Q, \Sigma, \Gamma, \delta, q_0, \text{accept, reject}), \]

where \( \Sigma = \{0, 1\} \), can be represented as a binary string as follows:

- \( Q = \{q_0, q_1, \ldots, q_n\} \), where \( q_{n-1} = \text{accept} \) and \( q_n = \text{reject} \)
- \( \Gamma = \{a_0, a_1, \ldots, a_m\} \), where \( a_0 = 0, a_1 = 1, a_2 = \square \)
- Let \( \Delta_0 = L \) and \( \Delta_1 = R \)
- The code for the transition function \( \delta \) rule

\[ \delta(q_i, a_j) = (q_r, a_s, \Delta_t) \]

is

\[ c_{ij} = 0^{i+1}10^{j+1}10^{r+1}10^{s+1}10^{t+1} \]

The code \( \langle M \rangle \) for the whole machine \( M \) is

\[ 111c_{00}11c_{01}11 \ldots 11c_{0m}11c_{10}11 \ldots 11c_{n-2,0}11 \ldots 11c_{n-2,m}111 \]

For example, the code for the following TM is

\[ \langle M \rangle = 1101010010100101001001100100100 \ldots 111 \]

\[ \delta(q_0, 0) = (q_0, \text{L}, R) \]
\[ \delta(q_0, 1) = (q_0, \text{L}, R) \]

Thus, every standard Turing machine \( M \) recognizing some language over the alphabet \( \{0, 1\} \) has a binary code \( \langle M \rangle \). On the other hand, we can associate some Turing machine \( M_b \) to each binary string \( b \).
A TM recognizing the language \( \{ 0^k | k \geq 0 \} \):

- \( q_0 \) \( \xrightarrow{0} 0, R \) \( q_1 \)
- \( \square \xrightarrow{\square}, L \)
- \( 1 \xrightarrow{1}, R \)
- \( q_2 \) \( q_3 \)

However, all binary strings are not codes for Turing machines. For instance, \( 00, 011110, 111000111 \) and \( 11101010111 \) are not legal codes for TMs according to the chosen encoding. We associate with illegal binary strings a trivial machine, \( M_{\text{triv}} \), rejecting all inputs:

\[
M_b = \begin{cases} 
M, & \text{if } b = \langle M \rangle \text{ is the code for Turing machine } M \\
M_{\text{triv}}, & \text{otherwise}
\end{cases}
\]
Hence, all Turing machines over \( \{0, 1\} \) can be enumerated:
\[ M_0, M_1, M_{00}, M_{01}, M_{10}, M_{11}, M_{000}, \ldots \]

At the same time we obtain an enumeration of the Turing-recognizable languages over \( \{0, 1\} \):
\[ L(M_0), L(M_1), L(M_{00}), L(M_{01}), L(M_{10}), \ldots \]

A language can appear more than once in this enumeration.

By diagonalization we can prove that the language \( D \) corresponding to the decision problem

*Does the Turing machine \( M \) reject its own description (\( \langle M \rangle \))?*

is not Turing-recognizable.

Hence, the decision problem corresponding to \( D \) is unsolvable.

**Lemma D** Language \( D = \{ b \in \{0, 1\}^* | b \notin L(M_b) \} \) is not Turing-recognizable.

**Proof.** Let us assume that \( D = L(M) \) for some standard Turing machine \( M \).

Let \( d = \langle M \rangle \) be the binary code of \( M \); i.e., \( D = L(M_d) \).

However,
\[ d \in D \iff d \notin L(M_d) = D. \]

We have a contradiction and the assumption cannot hold. Hence, there cannot exist a standard Turing machine \( M \) s.t. \( D = L(M) \).

Therefore, \( D \) cannot be Turing-recognizable. \( \square \)
4 Decidability

- The **acceptance problem** for DFAs: Does a particular DFA $B$ accept a given string $w$?
- Expressed as a formal language
  $$A_{DFA} = \{ B, w \mid B \text{ is a DFA that accepts string } w \}$$
- We simply need to represent a TM that decides $A_{DFA}$
- The Turing machine can easily simulate the operation of the DFA $B$ on input string $w$
- If the simulation ends up in an accept state, *accept*
- If it ends up in a nonaccepting state, *reject*
- DFAs can be represented (encoded) in the same vein as Turing machines

**Theorem 4.1** $A_{DFA}$ is a decidable language.

- NFAs can be converted to equivalent DFAs, and hence a corresponding theorem holds for them
- Regular expressions, on the other hand, can be converted to NFAs, and therefore a corresponding result also holds for them
- A different kind of a problem is *emptiness testing*: recognizing whether an automaton $A$ accepts any strings at all
  $$E_{DFA} = \{ A \mid A \text{ is a DFA and } L(A) = \emptyset \}$$

**Theorem 4.4** $E_{DFA}$ is a decidable language.
The TM can work as follows on input \( \langle A \rangle \):
1. Mark the start state of \( A \)
2. Repeat until no new states get marked:
   - Mark any state that has a transition coming in from any state that is already marked
3. If no accept state is marked, accept; otherwise, reject

Also determining whether two DFAs recognize the same language is decidable.

The corresponding language is \( EQ_{DFA} = \{ \langle A, B \rangle | A \text{ and } B \text{ are DFAs and } L(A) = L(B) \} \)

The decidability of \( EQ_{DFA} \) follows from Theorem 4.4 by turning to consider the symmetric difference of \( L(A) \) and \( L(B) \).

\[
L(C) = (L(A) \setminus L(B)) \cup (L(B) \setminus L(A)) = (L(A) \cap \overline{L(B)}) \cup (L(A) \setminus L(B))
\]

- \( L(C) = \emptyset \) if and only if \( L(A) = L(B) \)
- Because the class of regular languages is closed under complementation, union, and intersection, the TM can construct automaton \( C \) given \( A \) and \( B \)
- We can use Theorem 4.4 to test whether \( L(C) \) is empty

**Theorem 4.5** \( EQ_{DFA} \) is a decidable language.

Turning to context-free grammars we cannot go through all derivations because there may be an infinite number of them.
• In the acceptance problem we can consider the grammar converted into Chomsky normal form in which case any derivation of a string \( w, |w| = n \), has length \( 2n - 1 \).

• The emptiness testing of a context-free grammar \( G \) can be decided in a different manner:
  - Mark all terminal symbols in \( G \)
  - Repeat until no new variables get marked:
    - Mark any variable \( A \) where \( G \) has a rule \( A \rightarrow U_1 \ldots U_k \) and each symbol \( U_1 \ldots U_k \) has already been marked
    - If the start variable has not been marked, accept; otherwise reject
  - The equivalence problem for context-free grammars, on the other hand, is not decidable.

• Because of the decidability of the acceptance problem, all context-free languages are decidable.
• Hence, regular languages are a proper subset of context-free languages, which are decidable.
• Furthermore, decidable languages are a proper subset of Turing-recognizable languages.