Discrete Event Systems
Solution to Exercise Sheet 3

1 Pumping Lemma [Exam]

The Pumping Lemma in a Nutshell
Given a language $L$, assume for contradiction that $L$ is regular and has the pumping length $p$. Construct a suitable word $w \in L$ with $|w| \geq p$ ("there exists $w \in L$") and show that for all divisions of $w$ into three parts, $w = xyz$, with $|x| \geq 0$, $|y| \geq 1$, and $|xy| \leq p$, there exists a pumping exponent $i \geq 0$ such that $w' = xy^i z \notin L$. If this is the case, $L$ is not regular.

a) Language $L_1$ can be shown to be non-regular using the pumping lemma. Assume for contradiction that $L_1$ is regular and let $p$ be the corresponding pumping length. Choose $w$ to be the word $0110^p1$. Because $w$ is an element of $L_1$ and has length more than $p$, the pumping lemma guarantees that $w$ can be split into three parts, $w = xyz$, where $|xy| \leq p$ and for any $i \geq 0$, we have $xy^i z \in L_1$. In order to obtain the contradiction, we must prove that for every possible partition into three parts $w = xyz$ where $|xy| \leq p$, the word $w'$ cannot be pumped. We therefore consider the various cases.

1. If $y$ starts anywhere within the first three symbols (i.e. 011) of $w$, deleting $y$ (pumping with $i = 0$) creates a word with an illegal prefix (e.g. 1 0$^p$1 for $y = 01$).
2. If $y$ consists of only 0s from the second block, the word $w' = xy^2 z$ has more 0s than 1s in the last $|w'| − 3$ symbols and hence $c \neq d$.

Note that $y$ cannot contain 1s from the second block because of the requirement $|xy| \leq p$. We have shown that for all possible divisions of $w$ into three parts, the pumped word is not in $L_1$. Therefore, $L_1$ cannot be regular and we have a contradiction.

b) With the adapted language $L_2$, the proof of non-regularity is much more tricky! Specifically, non-regularity of $L_2$ cannot be proven using the pumping lemma, because any word in $L_2$ can actually be pumped! Consider for instance a word $w$ of the form $0110^p1^p$. In this case, we can split $w$ into the three parts $x = 0$, $y = 11$, $z = 0^p1^p$, which is in accordance with the rules of the pumping lemma. It can be seen, however, that any word $xy^i z$ is also in $L_2$! That is, the language $L_2$ can be pumped and yet, it is not regular as shown below.

Assume for contradiction that there exists a finite automaton $A$ which accepts the language $L_2$. Every word that starts with the input-sequence 0110 is only accepted if the remainder of the word has the form $0^c 1^d$ for some integer $c > 0$. Let $q_1$ be the state reached after the input 0110. Given the automaton $A$, we can construct a regular automaton $A'$ that is equivalent to $A$ with the only difference that its initial state is $q_1$. By the definition of $A$, this adapted finite automaton $A'$ accepts all words of the form $0^c 1^d$. However, as shown on slide 1/95 of the script, the language $0^c 1^d$ is not regular. Hence, $A'$ and thus $A$ cannot be finite automata. Because there exists a finite automaton for every regular language, it follows that $L_2$ cannot be regular. Language $L_2$ shows that while every regular language
can be pumped according to the pumping lemma, there are also non-regular languages that
be pumped.

Variant: We can alternatively use the fact that if two languages $L$ and $L'$ are regular, the
language defined by the intersection of the two languages $L \cap L'$ is regular as well (cf. p.
1/41). Consider the regular language $L_3 = \{ w \in 0110^*1^* \}$. Notice that the intersection of
$L_3$ with $L_2 = \{ 0^a1^b0^c1^d \mid a, b, c, d \geq 0 \text{ and if } a = 1 \text{ and } b = 2 \text{ then } c = d \}$ contains exactly
all words $w \in \{ 0110^*1^n \mid n \geq 0 \}$. This, however, is the exact language $L_1$ we proved not
to be regular in the first part of this exercise. If we assume $L_2$ to be regular, $L_1$ must be
regular as well, since $L_1 = L_2 \cap L_3$. This is a contradiction. Thus $L_2$ cannot be regular.

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| The argumentation above is based on the closure properties of regular languages and only
works in the direction presented. That is, for an operator $\phi \in \{ \cup, \cap, \cdot \}$, we have:

$$L_1 \text{ or } L_2 \text{ is non-regular, } \Rightarrow \text{ or } L_1 \text{ or } L_2 \text{ is regular.}$$

If either $L_1$ or $L_2$ or both are non-regular, we cannot deduce the non-regularity of $L$ or
vice-versa. Moreover, $L$ being regular does not imply that $L_1$ and $L_2$ are regular as well.
This may sound counter-intuitive which is why we give examples for the three operators.

- $L = L_1 \cup L_2$: Let $L_1$ be any non-regular language and $L_2$ its complement. Then
$L = \Sigma^*$ is regular.

- $L = L_1 \cap L_2$: Let $L_1$ be any non-regular language and $L_2$ its complement. Then $L = \emptyset$
is regular.

- $L = L_1 \cdot L_2$: Let $L_1 = \{ a^* \}$ (a regular language) and $L_2 = \{ a^p \mid p \text{ is prime} \}$ (a
non-regular language) then $L = \{ \text{aa}a^* \}$ is regular.

Hence, to prove that a language $L_\phi$ is non-regular, you assume it to be regular for contra-
diction. Then you combine it with a regular language $L_\phi$ to obtain a language $L = L_\phi \cdot L_\phi$.
If $L$ is non-regular, $L_\phi$ could not have been regular either.

2 Deterministic Finite Automata [Exam]

We could use the systematic transformation scheme presented in the lecture (slide 1/75). Con-
sidering the large number of states, however, this will easily lead to an explosion of states in
the derandomized automaton. Hence, we build the deterministic finite automaton in a step-wise
manner, only creating those states that are actually required: Initially, the automaton requires a
0. Subsequently, only a 1 is accepted. Including the various transitions, this 1 can lead to three
different states, namely states 2, 3, and 4.

![Diagram of automaton](image)

In any of the states 2, 3, and 4, only a 1 is accepted. Assume that the automaton is currently
in state 2, this 1 can lead to states $\{2, 3, 4\}$ when including all $\varepsilon$-transitions. When in state 3,
the 1 leads to states $\{2, 3, 4, 5\}$ and finally, when being in state 4, the reachable states given
a 1 are $\{2, 3, 4\}$. Hence, a 1 leads from state $\{2, 3, 4\}$ to state $\{2, 3, 4, 5\}$. Repeating the same
process for state $\{2, 3, 4, 5\}$, we can see that, again, only a 1 is accepted, which leads to state
$\{2, 3, 4, 5, 6\}$. Because the state 6 in the original NFA was an accepting state, $\{2, 3, 4, 5, 6\}$ is also
accepting in the DFA. From state $\{2, 3, 4, 5, 6\}$, an additional 1 will lead to another accepting
state $\{1, 2, 3, 4, 5, 6\}$. And from this state, any subsequent 1 returns to state $\{1, 2, 3, 4, 5, 6\}$ as
well.
What happens if a 0 occurs in the input? This is feasible only when the deterministic state includes either state 1 or state 6. In state \( \{2, 3, 4, 5, 6\} \), a 0 necessarily leads to state \( \{4\} \), whereas in state \( \{1, 2, 3, 4, 5, 6\} \) a 0 leads to state \( \{2, 4\} \). In both of these states, the only acceptable input symbol is a 1 and leads to the state \( \{2, 3, 4\} \). Hence, the deterministic finite automaton looks like this:

![Automaton Diagram]

It can easily be seen, that first the states \( \{4\} \), \( \{2, 4\} \) and then the states \( \{2, 3, 4, 5, 6\}, \{1, 2, 3, 4, 5, 6\} \) can be merged and hence, the automaton can be reduced to the one shown in the next figure.

![Automaton Diagram]

This is not a DFA yet, because the crash state is still missing. The final deterministic automaton looks like this:

![Automaton Diagram]

### 3 Transforming Automata [Exam]

The regular expression can be obtained from the finite automaton using the transformation presented in the script on slide 1/85. After ripping out state \( q_2 \), the corresponding GNFA looks like this:

![GNFA Diagram]
After also removing state $q_1$, the GNFA looks as follows.

$$0 \cup 11^*0(01^*0)^*1$$

![Diagram of GNFA](image)

Eliminating the last state $q_3$ yields the final solution, which is $(01^*0)^*1(0 \cup 11^*0(01^*0)^*1)^*$.

Note: Ripping out the interior states in a different order yields a distinct yet equivalent regular expression. The order $q_3, q_2, q_1$, for example, results in $((0 \cup 10^*1)^*0\cup 1)^*$.

4 Regular and Context-Free Languages

a) Sometimes, even simple grammars can produce tricky languages. We can interpret the $1$s and $2$s of the second production rule as opening and closing brackets. Hence, $L(G)$ consists of all correct bracket terms where at least one $0$ must be in each bracket.

Choose $w = 1^p02^p \in L(G)$. Let $w = xyz$ with $|xy| \leq p$ and $|y| \geq 1$ (pumping lemma). Because of $|xy| \leq p$, $xy$ can only consist of $1$s. According to the pumping lemma, we should have $xy^iz \in L$ for all $i \geq 0$. However, by choosing $i = 0$ we delete at least one $1$ and get a word $w' = 1^{p-|y|}02^p$ with $|y| \geq 1$. $w'$ is not in $L$ since it has fewer $1$s than $2$s. This means that $w$ is not pumpable and hence, $L(G)$ is not regular.

b) Since every regular language is also context-free, we can choose an arbitrary regular language. For example, we can choose the language $L = \{0^n1, n \geq 1\}$ which is clearly regular. A context-free grammar for this language uses only the production $S \rightarrow 0S \mid 1$.

5 Context-Free Grammars

a) An example for a grammar $G$ producing the language $L_1$ is $G = (V, \Sigma, R, S)$ with

$V = \{X, A\}$,

$\Sigma = \{0, 1\}$,

$R = \{X \rightarrow XAX \mid A\}$,

$S = X$

Note: The language is regular!

b) A rather natural grammar generating $L_2$ uses the following productions:

$$S \rightarrow A1A$$

$$A \rightarrow A1 \mid 1A \mid A01 \mid 0A1 \mid 01A \mid A10 \mid 1A0 \mid 10A \mid \varepsilon$$

Another slightly more complicated solution yielding simpler productions looks as follows:

$$S \rightarrow A1A$$

$$A \rightarrow AA \mid 1A0 \mid 0A1 \mid 1 \mid \varepsilon$$

The idea of both grammars is to first ensure that there is at least one $1$ more and then have a production that generates all possible strings with the same number of $0$s and $1$s or further $1$s at arbitrary places.
6 Pushdown Automata

a) \( \epsilon, 0, 00, (0), 0(0), (0), 00 \)

b) It is ambiguous, because the word 00 has two different leftmost derivations.

\[
\begin{align*}
S & \rightarrow SA & S & \rightarrow SA \\
& \rightarrow A & \rightarrow SAA \\
& \rightarrow AA & \rightarrow AA \\
& \rightarrow 0A & \rightarrow 0A \\
& \rightarrow 00 & \rightarrow 00 
\end{align*}
\]

It can also be seen by taking a look at these two derivation trees that both belong to the word 00:

Because the two derivation trees are structurewise different, the word 00 can be derived ambiguously from \( G \).

**Ambiguity of Grammars**

*Definition:* A string \( s \) is derived *ambiguously* in a context-free grammar \( G \) if it has two or more different leftmost/rightmost derivations (or two structurewise different derivation trees). Grammar \( G \) is *ambiguous* if it generates some string ambiguously. A leftmost/rightmost derivation replaces in every step the leftmost/rightmost variable.

*Example:* The grammar with the productions ‘\( S \rightarrow S \cdot S \mid S + S \mid a \)’ is ambiguous since the string \( s = a \cdot a + a \) has two different leftmost derivations.

\[
\begin{align*}
S & \rightarrow S \cdot S & S & \rightarrow S + S \\
& \rightarrow a \cdot S & \rightarrow S \cdot S + S \\
& \rightarrow a \cdot S + S & \rightarrow a \cdot S + S \\
& \rightarrow a \cdot a + S & \rightarrow a \cdot a + S \\
& \rightarrow a \cdot a + a & \rightarrow a \cdot a + a 
\end{align*}
\]

Intuitively, the derivation on the left corresponds to the arithmetic expression \( a \cdot (a + a) \) because we first derive a product and then substitute one factor by a sum while the derivation on the right corresponds to \( (a \cdot a) + a \) because we first have a sum and then substitute one summand by a product.

The productions of an equivalent non-ambiguous grammar are \( A \rightarrow S + a \mid S \cdot a \mid a \).
A simple non-deterministic PDA for $L(G)$ looks as follows:

\[
\begin{align*}
\), (\rightarrow & \varepsilon \\
\), \varepsilon & \rightarrow ( \\
0, \varepsilon & \rightarrow \varepsilon \\
\end{align*}
\]

\[
\begin{array}{c}
\text{ε, ε \rightarrow $} \\
\text{$ \rightarrow \varepsilon}
\end{array}
\]

\[
\begin{array}{c}
\text{ε, § \rightarrow ε}
\end{array}
\]

Deterministic PDAs

A push-down automaton $M$ is deterministic if in each state, there is exactly one successor state for every combination $(a, b) \in \Sigma \times \Gamma$ where $\Sigma$ is the string input alphabet and $\Gamma$ is the stack alphabet. Note that if a state $q$ has only one outgoing transition `$, ε \rightarrow $' the PDA is still deterministic since there is no ambiguity of what the successor state of $q$ will be. If a state $q$, however, has two outgoing transitions, `$, ε \rightarrow x$' and `$, b \rightarrow y$' leading into different states, it is unclear which transition the system should take if the string input in state $q$ is `$a'$ and the top element on the stack is `$b'$. A PDA containing such ambiguous transitions is not deterministic.

Unlike in deterministic finite automata, we take the liberty of omitting transitions leading to an (imaginary) fail state as well as the fail state itself when drawing deterministic PDAs.

Considering this, the PDA given above is not deterministic: From the middle state, there are two transitions `(, ε \rightarrow ( and `, § \rightarrow ε', such that we do not know which one to take if we read a `( while the top element on the stack is `§'. We can overcome this problem in different ways.

If we assume that our PDA recognizes the end of the input string (denoted by `−'), it is easy to transform the non-deterministic PDA above into a deterministic one:

\[
\begin{align*}
\), (\rightarrow & \varepsilon \\
\), \varepsilon & \rightarrow ( \\
0, \varepsilon & \rightarrow \varepsilon \\
\end{align*}
\]

\[
\begin{array}{c}
\text{ε, ε \rightarrow $} \\
\text{−, § \rightarrow ε}
\end{array}
\]

If we assume that the PDA is not able to determine the end of the input, it is not that easy to derive the deterministic PDA from the non-deterministic one.

An example of a deterministic PDA accepting $L(G)$ is the following:

\[
\begin{align*}
\), (\rightarrow & \varepsilon \\
0, \varepsilon & \rightarrow \varepsilon \\
\end{align*}
\]

\[
\begin{array}{c}
\text{ε, ε \rightarrow $} \\
\text{ε, § \rightarrow §}
\end{array}
\]
The deterministic PDA using as few states as possible is the following: