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Lower Bounds for Synchronizing Word Lengths in Partial Automata

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It was conjectured by Černý in 1964, that a synchronizing DFA on n states always has a synchronizing word of length at most $(n - 1)^2$, and he gave a sequence of DFAs for which this bound is reached. Until now a full analysis of all DFAs reaching this bound was only given for $n \leq 5$, and with bounds on the number of symbols for $n \leq 12$. Here we give the full analysis for $n \leq 7$, without bounds on the number of symbols.

For PFAs (partial automata) on ≤ 7 states we do a similar analysis as for DFAs and find the maximal shortest synchronizing word lengths, exceeding $(n - 1)^2$ for $n \geq 4$. Where DFAs with long synchronization typically have very few symbols, for PFAs we observe that more symbols may increase the synchronizing word length. For PFAs on ≤ 10 states and two symbols we investigate all occurring synchronizing word lengths.

We give series of PFAs on two and three symbols, reaching the maximal possible length for some small values of n . For $n = 6, 7, 8, 9$, the construction on two symbols is the unique one reaching the maximal length. For both series the growth is faster than $(n - 1)^2$, although still quadratic.

Based on string rewriting, for arbitrary size we construct a PFA on three symbols with exponential shortest synchronizing word length, giving significantly better bounds than earlier exponential constructions. We give a transformation of this PFA to a PFA on two symbols keeping exponential shortest synchronizing word length, yielding a better bound than applying a similar known transformation. Both PFAs are transitive.

Finally, we show that exponential lengths are even possible with just one single undefined transition, again with transitive constructions.

Keywords: DFA, PFA, careful synchronization, Černý conjecture

1. Introduction and Preliminaries

A *deterministic finite automaton (DFA)* over a finite alphabet Σ is called *synchronizing*, if it admits a *synchronizing word*. A word $w \in \Sigma^*$ is called *synchronizing* (or directed, or reset), if, starting in any state q , after reading w , one always ends in one

particular state q_s . So reading w acts as a reset button: no matter in which state the system is, it always moves to the particular state q_s . Now Černý's conjecture [2] states:

Every synchronizing DFA on n states admits a synchronizing word of length $\leq (n - 1)^2$.

Surprisingly, despite extensive effort, this conjecture is still open, and even the best known upper bounds are still cubic in n . In 1983 Pin [17] established the bound $\frac{1}{6}(n^3 - n) - 1$ for $n > 4$, based on [9] and Pin's thesis. Recently, a slight asymptotic improvement to Pin's bound has been obtained by Szykuła [20] (effective for $n \geq 724$). For a survey on synchronizing automata and Černý's conjecture, we refer to [22].

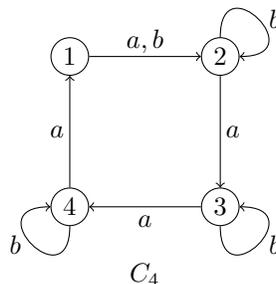
Formally, a *deterministic finite automaton (DFA)* over a finite alphabet Σ consists of a finite set Q of states and a map $\delta : Q \times \Sigma \rightarrow Q$.^a For $w \in \Sigma^*$ and $q \in Q$, we define qw inductively by $q\lambda = q$ and $qaw = \delta(q, a)w$ for $a \in \Sigma$, where λ is the empty word. So qw is the state where one ends, when starting in q and reading the symbols in w consecutively, and qa is a short hand notation for $\delta(q, a)$. A word $w \in \Sigma^*$ is called *synchronizing*, if a state $q_s \in Q$ exists such that $qw = q_s$ for all $q \in Q$.

In [2], Černý already gave DFAs for which the bound of the conjecture is attained: for $n \geq 2$ the DFA C_n is defined to consist of n states $1, 2, \dots, n$, and two symbols a, b , acting by $qa = q + 1$ for $q = 1, \dots, n - 1$, $\delta(n, a) = 1$, and $qb = q$ for $q = 2, \dots, n$, $1b = 2$.

For $n = 4$, this is depicted on the right. For C_n , the string $w = b(a^{n-1}b)^{n-2}$ of length $|w| = (n - 1)^2$ satisfies $qw = 2$ for all $q \in Q$, so w is synchronizing. No shorter synchronizing word exists for C_n , as is shown in [2], showing that the bound in Černý's conjecture is sharp.

One goal of this paper is to investigate the synchronizing word lengths of all DFAs on at most 7 states. We also search for the maximal word lengths when restricting to DFAs with a given alphabet size. The main result on DFAs is that Černý's conjecture is true for $n \leq 7$. Our results extend those in [12], in which Černý's conjecture is verified for $n \leq 5$. A complete analysis of all DFAs of $n = 6$ and $n = 7$ states is not provided in [12]: the number of symbols is limited to 6 and 4 respectively. The computations in [12] extend several results by the same authors.

A generalization of a DFA is a *Partial Finite Automaton (PFA)*; the only difference is that now the transition function δ is allowed to be partial. In a PFA, qw may be undefined, in fact it is only defined if every step is defined. A word



^aFor synchronization the initial state and the set of final states in the standard definition may be ignored.

$w \in \Sigma^*$ is called *carefully synchronizing* for a PFA, if a state $q_s \in Q$ exists such that qw is defined and $qw = q_s$ for all $q \in Q$. In other words: starting in any state q and reading w , every step is defined and one always ends in state q_s . A PFA, in particular a DFA, is called *transitive* or *strongly connected* if for every ordered pair (q, q') of states, there is a word $w \in \Sigma^*$ such that $qw = q'$. As being a generalization of DFAs, the shortest carefully synchronizing word may be longer. For all $n \geq 4$ we show that this is indeed the case: for $n = 4, 5, 6, 7$ we find the maximal shortest carefully synchronizing word length to be 10, 21, 37 and 63, respectively.

Also for PFAs we investigate the dependence on the alphabet size. To exclude infinitely many trivial extensions, we only consider *basic* PFAs: no two symbols act in the same way, no symbol acts as the identity and no symbol is a restricted version of either another symbol or the identity. Obviously, these properties have no influence on synchronization. Somewhat surprisingly, we find that larger alphabets may lead to longer carefully synchronizing words, in contrast to the situation for DFAs.

We compute all binary PFAs with up to 10 states, to obtain all possible synchronization lengths, both for DFAs and proper PFAs. For DFAs, the authors of [12] got as far as 12 states with obtaining these lengths. With that, they extended the maximum number of states from earlier analyses, by themselves and by others, ranging from 9 states to 11 states. The authors of [12] obtained all possible synchronization lengths for ternary DFAs with 8 states as well. Several gaps exist in the ranges of synchronization lengths for binary DFAs. It appears that such gaps also exist for binary PFAs.

For every n we give a PFA on n states and 2 symbols for which we exactly compute the shortest carefully synchronizing word length, for every $n \geq 6$ strictly exceeding $(n - 1)^2$. This length is quadratic in n , but it is not a polynomial: the precise formula deals with Fibonacci numbers. For $n = 6, 7, 8, 9$ this is the only construction giving the maximal shortest synchronizing word length for binary PFAs. Similarly, we give a sequence of PFAs on three symbols, reaching the maximal length for $n = 3, 4, 5$.

For PFAs the maximal length grows exponentially in n , as was already observed by Rystsov [19]. Rystsov established the lower bound $\Omega((3 - \varepsilon)^{n/3})$ and the upper bound $O((3 + \varepsilon)^{n/3})$. The upper bound can be found in [10] as well. Martyugin [14] established the lower bound $\Omega(3^{n/3})$ with a construction in which the number of symbols is linear in n .

In [13], the author Martyugin obtained a lower bound for the synchronization of PFAs with a constant alphabet size, which lies between polynomial and exponential, as a result of an elegant construction of a series of PFAs (see also the last section of [4]). In [16], the same author obtained a near-exponential lower bound, using a different construction of PFAs. In [23] it was shown that exponential bounds exist for every constant alphabet size being at least two. For two symbols the bound $\Omega(2^{n/36})$ was given for the transitive case and the bound $\Omega(2^{n/26})$ for the general case. Our construction strongly improves this and gives length $\Omega(2^{n/3 - 3 \log_2(n)/2}) =$

$\Omega(2^{n/3}/n^{3/2})$ for binary PFAs and length $\Omega(2^{2n/5-\log_2(n)}) = \Omega(2^{2n/5}/n)$ for ternary PFAs, with transitive constructions. Finally we show that both in the binary and ternary case exponential growth is even possible with a single undefined transition, again with transitive constructions. For the ternary construction, the key idea is that synchronization is forced to mimic exponentially many string rewrite steps, similar to binary counting. This ternary PFA can be transformed to a binary PFA by a standard construction for which we develop a substantial improvement.

The decision problems which correspond to our asymptotic constructions are PSPACE-complete, if we do not take transitivity into account. This follows from [15], in which the most specific decision problem is treated, namely the problem of determining if a binary PFA with only one undefined transition is carefully synchronizing. The fact that this problem is PSPACE-complete already suggested the existence of a nonpolynomial construction, because otherwise we would have had PSPACE = NP. However, the construction in [15] is not transitive. Using [23, Lemma 2] and [23, Lemma 6], one can make the construction transitive, but the property of having only one undefined transition will be affected. So if we do take transitivity into account, then PSPACE-completeness is obtained for the decision problems which correspond to our asymptotic constructions, except the last one (with only one undefined transition).

The basic tool to analyze (careful) synchronization is the *power automaton*. For any DFA or PFA (Q, Σ, δ) , its power automaton is the DFA $(2^Q, \Sigma, \delta')$ where $\delta' : 2^Q \times \Sigma \rightarrow 2^Q$ is defined by $\delta'(V, a) = \{q \in Q \mid \exists p \in V : \delta(p, a) = q\}$, if $\delta(p, a)$ is defined for all $p \in V$, otherwise $\delta'(V, a) = \emptyset$. For any $V \subseteq Q, w \in \Sigma^*$, we define Vw as before, using δ' instead of δ . From this definition, one easily proves that $Vw = \{qw \mid q \in V\}$ if qw is defined for all $q \in V$, otherwise $Vw = \emptyset$, for any $V \subseteq Q, w \in \Sigma^*$. A set of the shape $\{q\}$ for $q \in Q$ is called a *singleton*. So a word w is (carefully) synchronizing, if and only if Qw is a singleton. Hence a DFA (PFA) is (carefully) synchronizing, if and only if its power automaton admits a path from Q to a singleton, and the shortest length of such a path corresponds to the shortest length of a (carefully) synchronizing word.

This paper is an extended version of the DLT2017 paper [5]. It contains several new contributions, in particular:

- Sec. 2 – For DFAs we extend the complete analysis from $n \leq 6$ to $n \leq 7$.
 - We further investigate DFAs with given alphabet size.
- Sec. 3 – For PFAs we also extend the analysis to $n \leq 7$, and fine tune it by also taking the number of symbols into account.
 - We investigate the carefully synchronizing word lengths for binary PFAs on $n \leq 10$ states.
- Sec. 4 – We give sequences of binary and ternary PFAs, reaching the maximal possible length for some values of n .
- Sec. 7 – We improve our asymptotic results and include a construction with a single undefined transition.

The most important update to Sections 5 and 6 is that transitivity is taken into account. Section 5 presents our construction of PFAs on three symbols with exponential shortest carefully synchronizing word length. In Section 6 we improve the transformation used by Martyugin [14] and Vorel [23] to reduce to alphabet size two. We conclude in Section 8.

2. Critical DFAs on at Most 7 States

A natural question when studying Černý's conjecture is: what can be said about automata in which the bound of the conjecture is actually attained, the so-called critical automata? Throughout this section we restrict ourselves to basic DFAs. As has already been noted by several authors [7, 21, 22], critical DFAs are rare. There is only one construction known which gives a critical DFA for each n , namely the well-known sequence C_n , discovered by and named after Černý [2]. Apart from this sequence, all known critical DFAs have at most 6 states. In [7], all critical DFAs on less than 5 states were identified, without restriction on the size of the alphabet. For $n = 5$ and $n = 6$ it was still an open question if there exist critical (or even supercritical) DFAs, other than those already discovered by Černý, Roman [18] and Kari [11]. In [5], we verified that this is not the case, so for $n = 5$ only two critical DFAs exist (Černý, Roman) and also for $n = 6$ only two exist (Černý, Kari). Here we extend the analysis to $n = 7$, for which Černý's DFA is the only critical DFA. In fact our results also prove the following theorem:

Theorem 1. *Every synchronizing DFA on $n \leq 7$ states admits a synchronizing word of length at most $(n - 1)^2$.*

As Trahtman already noted in his paper [21], for $n \geq 6$ there seems to be a gap in the range of possible shortest synchronization lengths. For example, his analysis showed that there are no DFAs on 6 states with shortest synchronizing word length 24, and no DFAs on 7 states with length 33, 34 or 35, when restricting to at most 4 symbols. Our analysis shows that this is true without restriction on the alphabet: there is no DFA on 6 states with shortest synchronizing word length 24. For $n \leq 6$ all other lengths are feasible: if $n \leq 6$ and $1 \leq k \leq (n - 1)^2$, $k \neq 24$, then there exists a DFA on n states with shortest synchronizing word length exactly k . For $n = 7$ all values $k \leq 32$ occur as shortest synchronizing word length.

As the number of DFAs on n states grows like 2^{n^n} , an exhaustive search is a non-trivial affair, even for small values of n . The problem is that the alphabet size in a basic DFA can be as large as $n^n - 1$. Up to now, for $n = 5, 6, 7$ only DFAs with at most four symbols were checked by Trahtman [21]. Here we describe our algorithm to investigate all DFAs on 5, 6 and 7 states, without restriction on the alphabet size.

Before explaining the algorithm, we introduce some terminology. A DFA \mathcal{B} obtained by adding some symbols to a DFA \mathcal{A} will be called an *extension* of \mathcal{A} . If $\mathcal{A} = (Q, \Sigma, \delta)$, then $S \subseteq Q$ will be called *reachable* if there exists a word $w \in \Sigma^*$

such that $Qw = S$. We say that S is *reducible* if there exists a word w such that $|Sw| < |S|$, and we call w a *reduction word* for S . Our algorithm is mainly based on the following observation:

Property 1. *If a DFA \mathcal{A} is synchronizing, and \mathcal{B} is an extension of \mathcal{A} , then \mathcal{B} is synchronizing as well and its shortest synchronizing word is at most as long as the shortest synchronizing word for \mathcal{A} .*

The algorithm roughly runs as follows. We search for (super)critical DFAs on n states, so a DFA is discarded if it synchronizes faster, or if it does not synchronize at all. For a given DFA $\mathcal{A} = (Q, \Sigma, \delta)$ which is not yet discarded or investigated, the algorithm does the following:

- (1) If \mathcal{A} is synchronizing and (super)critical, we have identified an example we are searching for.
- (2) If \mathcal{A} is synchronizing and subcritical, it is discarded, together with all its possible extensions (justified by Property 1).
- (3) If \mathcal{A} is not synchronizing, then find an upper bound L for how fast any synchronizing extension of \mathcal{A} will synchronize (see below). If $L < (n - 1)^2$, then discard \mathcal{A} and all its extensions. Otherwise, discard only \mathcal{A} itself.

The upper bound L for how fast any synchronizing extension of \mathcal{A} will synchronize, is found by analyzing distances in the directed graph of the power automaton of \mathcal{A} . For $S, T \subseteq Q$, the distance $\text{dist}(S, T)$ from S to T in this graph is equal to the length of the shortest word w for which $Sw = T$, if such a word exists.

The distances in the directed graph of the power automaton are computed by way of the Floyd-Warshall algorithm. As the computation complexity of the Floyd-Warshall algorithm is cubic, the complexity in terms of n is $\Theta(8^n)$, which is actually quite bad. For that reason, we took the effort to implement it far more efficiently than the straightforward way, see [3].

We do not compute $\text{dist}(S, T)$ if T is a singleton. Instead, we compute

$$\min\{\text{dist}(S, T) \mid T \subset Q \text{ and } |T| \leq i\}$$

for every $S \subseteq Q$ and $i = 1, 2, \dots, n - 1$: for $i = 1$ as a replacement, yielding vacated space in the distance matrix, and for larger i as a usage of this space.

A possible upper bound L is as follows:

- (1) Determine the size $|S|$ of a smallest reachable set S . Let m be the minimal distance from Q to a set of size $|S|$.
- (2) For each $k \leq |S|$, partition the collection of irreducible sets of size k into strongly connected components. Let m_k be the number of components plus the sum of their diameters.
- (3) For each reducible set R of size $k \leq |S|$, find the length l_R of its shortest reduction word. Let l_k be the maximum of these lengths.

- (4) Now note that a synchronizing extension of \mathcal{A} will have a synchronizing word of length at most

$$L = \sum_{k=2}^{|S|} (m_k + l_k) + m.$$

A slightly better upper bound is the following. Let M be the maximum distance from Q to a set of size $|S|$. Partition the irreducible sets of size $|S|$ which can be reached from Q into strongly connected components, and let c be the number of components plus the sum of their diameters. Then a synchronizing extension of \mathcal{A} will have a synchronizing word of length at most

$$L' = \sum_{k=2}^{|S|} (m_k + l_k) - c + 1 + M.$$

So one can say that Q as a reducible subset is treated differently in the construction of L' than in the construction L . As a consequence, $L' \leq L$, so L' is a better upper bound than L . In the upper bound L'' which is actually used in the computations, we extend this different treatment to other reducible subsets.

But first, we describe L in an inductive way. We take $L = L_{|S|} + m$, and define

$$\begin{aligned} L_1 &= 0, \\ L_k &= m_k + l_k + L_{k-1} \\ &= m_k + \max\{l_R \mid R \text{ is reducible and } |R| = k\} + L_{k-1} \quad \text{if } k > 1. \end{aligned}$$

Here, L_k is an upper bound for the maximum length of the shortest synchronizing word for any subset of size k . We take $L'' = L''_Q$, and we define inductively an upper bound L''_R for the length of the the shortest synchronizing word for a reducible subset R , and an upper bound L''_k for the maximum length of the shortest synchronizing word for any subset of size k . Define S_R , m_R , M_R and c_R as S , m , M and c respectively, but with Q replaced by R .

$$\begin{aligned} L''_R &= m_R && \text{if } |S_R| = 1, \\ L''_R &= L''_{|S_R|} - c_R + 1 + M_R && \text{if } |S_R| > 1, \\ L''_1 &= 0, \\ L''_k &= m_k + \max\{L''_{k-1}, L''_R \mid R \text{ is reducible and } |R| = k\} && \text{if } k > 1. \end{aligned}$$

Although L'' yields a better upper bound than L' in general, we do not always have $L'' \leq L'$. To overcome this, we improved the definition of L''_R in the newest version of the code, but only for $R \neq Q$, by taking the minimum of what is given above, and $L''_{|R|-1} + l_R$. (The calculations on DFAs with 7 states have not been redone.)

The algorithm performs a depth-first search. So after investigating a DFA, first all its extensions (not yet considered) are investigated before moving on. Still, we can choose which extension to pick first. We would like to choose an extension that is likely to be discarded immediately together with all its extensions. Therefore, we

apply the following heuristic: for each possible extension \mathcal{B} by one symbol, we count how many pairs of states in \mathcal{B} would be reducible. The extension for which this is maximal is investigated first. The motivation is that a DFA is synchronizing if and only if each pair is reducible [2].

Furthermore, we only investigate extensions \mathcal{B} by one symbol if either the number of pairs which synchronize in \mathcal{B} is larger than in \mathcal{A} , or \mathcal{A} (and hence also \mathcal{B}) is synchronizing. The idea behind this is the following, which is easy to prove. If \mathcal{A} is not synchronizing and \mathcal{B} is an extension of \mathcal{A} which is synchronizing, then \mathcal{B} has a symbol, which, when added to \mathcal{A} , increases the number of synchronizing pairs.

The algorithm which has actually been used also takes symmetries on the set of states into account, making it almost $n!$ times faster. The symmetry reduction on the states is perfect for automata which do not have a pair of conjugate symbols (two symbols a and b are conjugate if there exists a symmetry σ such that $\sigma b \sigma^{-1} = a$). Furthermore, we used a multithreaded version of the algorithm for the case of $n = 7$ states.

In the table below, we counted for every number of symbols (alph. size) and every minimal synchronization length (sync.) ≥ 31 , the number of corresponding basic DFAs with seven states, up to symmetry. We do not require the automata to be minimal, meaning that we allow solutions from which symbols can be removed without changing the synchronization length. This explains why our numbers differ from those found by Szykuła in his thesis.

alph. size	sync. 36	sync. 35	sync. 34	sync. 33	sync. 32	sync. 31
1						
2	1				3	3
3					3	8
4						4
total	1	0	0	0	6	15

Tables for less than seven states can be found in [6], which is an extended version of [7]. See also the graph in subsection 3.2.

In [12], the synchronization upper bounds which are used for pruning the search are different, and only work for DFAs. But there are also differences in the way the searching is performed. In [12], the searching is done by way of breadth-first search instead of depth-first search, taking far less overhead and resources, so it can be done for larger number of states as well. But it leads to more redundancy in the search: more DFAs need to be scanned, increasing the computation time. It appeared that with our search algorithm, one can compute all critical DFAs with 6 states within a day (using [12, Theorem 1] as a synchronization upper bound for pruning; with our synchronization upper bound it is about 400 times faster.)

3. PFAs with Small State Set

In the remainder of this paper, we study PFAs and shortest carefully synchronizing word lengths. In this section and the next, we focus on PFAs that have long shortest synchronizing words and a small number of states. In later sections we construct PFAs on two or three symbols with shortest carefully synchronizing words of exponential length for general n .

3.1. PFAs on at Most 7 States

To find PFAs with small number of states and long shortest carefully synchronizing word, we exploit that Property 1 also holds for PFAs. However, for PFAs it is not true that reducibility of all pairs of states guarantees careful synchronization. Therefore, we apply a different search algorithm. We search for a PFA with synchronizing length equal to or greater than some given target length. To construct it, we build the alphabet by choosing the symbols of a long shortest synchronizing word from left to right. More precisely, on the stack of the search function we always have a prefix of a possible synchronizing word. The search is pruned in the following three cases, where w is the prefix on the stack:

- (1) There exists a word u consisting of the letters of w , with $|u| < |w|$, such that either $Qu = Qw$, or Qu and Qw are both singletons;
- (2) The automaton \mathcal{A} , whose symbols are the letters of w , has a synchronizing word which is shorter than the target length;
- (3) The value of the upper bound L'' for the automaton \mathcal{A} is smaller than the target length.

If the search is not pruned, the prefix w will be extended by one letter a . To reduce the number of solutions and speed up the algorithm even further, we only select a candidate symbol a as follows:

- (1) If $Qwa = Qwb$ for a letter b of w , then a is only selected if it is equal to the first such letter in w ;
- (2) If $Qwa = Qwb$ does not hold for any letter b of w , then a is only selected if it is undefined outside Qw .

The purpose of symbol a is to get from Qw to the next subset. In the situation of (1), no new symbol need to be added to make the transition from Qw to the next subset. We choose a to be an old symbol, because there is no need to add a new symbol at this point in the search. In the situation of (2), we choose a to be defined on states of Qw only, because the purpose of a is to get from Qw to the next subset. There is no need to add a more complete symbol at this point in the search.

The selection rules (1) and (2) above significantly reduce the number of cases, but (2) has the drawback that the algorithm does not necessarily find the solution with the smallest possible alphabet any more. For example, it did not find a solution

of length 37 with only 6 symbols for $n = 6$. But postprocessing all solutions for $n = 6$ did reveal a solution of length 37 with only 6 symbols indeed.

During the postprocessing of a solution, symbols are made more complete, so (2) does not hold any longer. There are many ways to make the symbols more complete, but most of them will affect the synchronization length, which gives us effective pruning. For every solution with more complete symbols, symbols may have become the same, and we count the number of distinct symbols.

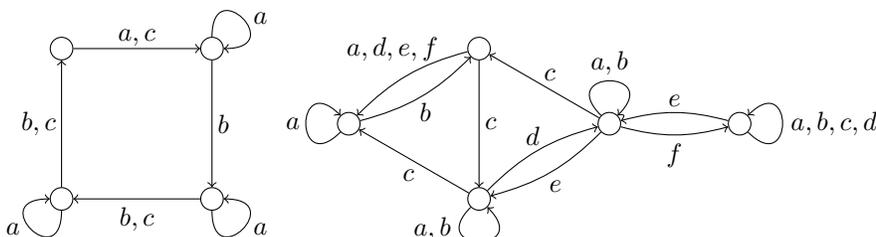
Just as for the DFAs, we took symmetry into account. But we did not need a multithreaded version of the algorithm for the case of $n = 7$ states.

For $n \leq 7$, our algorithm has identified the maximal length $p(n)$ of a shortest carefully synchronizing word in a PFA on n states. The results are:

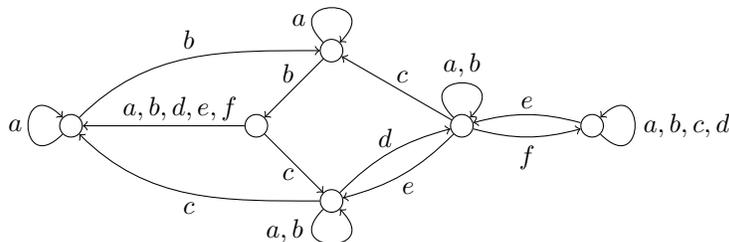
n	2	3	4	5	6	7
$p(n)$	1	4	10	21	37	63

For $n = 8$ states, 102 can be reached as shortest carefully synchronizing word length, using 9 symbols. But 8 states are too many for us to prove computationally that this is the largest possible length.

Whereas for $n \geq 6$, no critical DFAs are known with more than two symbols, PFAs with long shortest carefully synchronizing word lengths tend to have more symbols: for $n = 4, 5, 6, 7$ states, the minimal numbers of symbols achieving the maximal shortest carefully synchronizing word lengths 10, 21, 37 and 63 are 3, 6, 6 and 8 respectively. Below we give examples of PFAs on 4, 5, 6 and 7 states reaching these lengths.

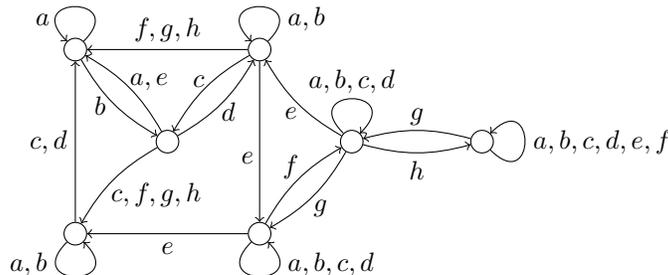


The left one has two synchronizing words of length 10: $abcabab(b+c)ca$. The right one has unique shortest synchronizing word $abcabdbecabdbfbcdca$ of length 21.



The shortest synchronizing word is $ab^2ab^2cb^2ab^2db^2eb^2cb^2ab^2db^2fb^2cdcb^2a$ for this

PFA on 6 states. It is unique and has length 37.



There are 81 shortest synchronizing words (of length 63) for this PFA on 7 states, all being of the form

$$abcabdbecabdbfdbgbdbbecabdbfdbhbdbeb.....bdefgeca.$$

This word is remarkably similar to the one for 5 states and also the actions of some of the symbols are comparable. It is however not yet sufficient to detect a pattern that could be extrapolated to larger n .

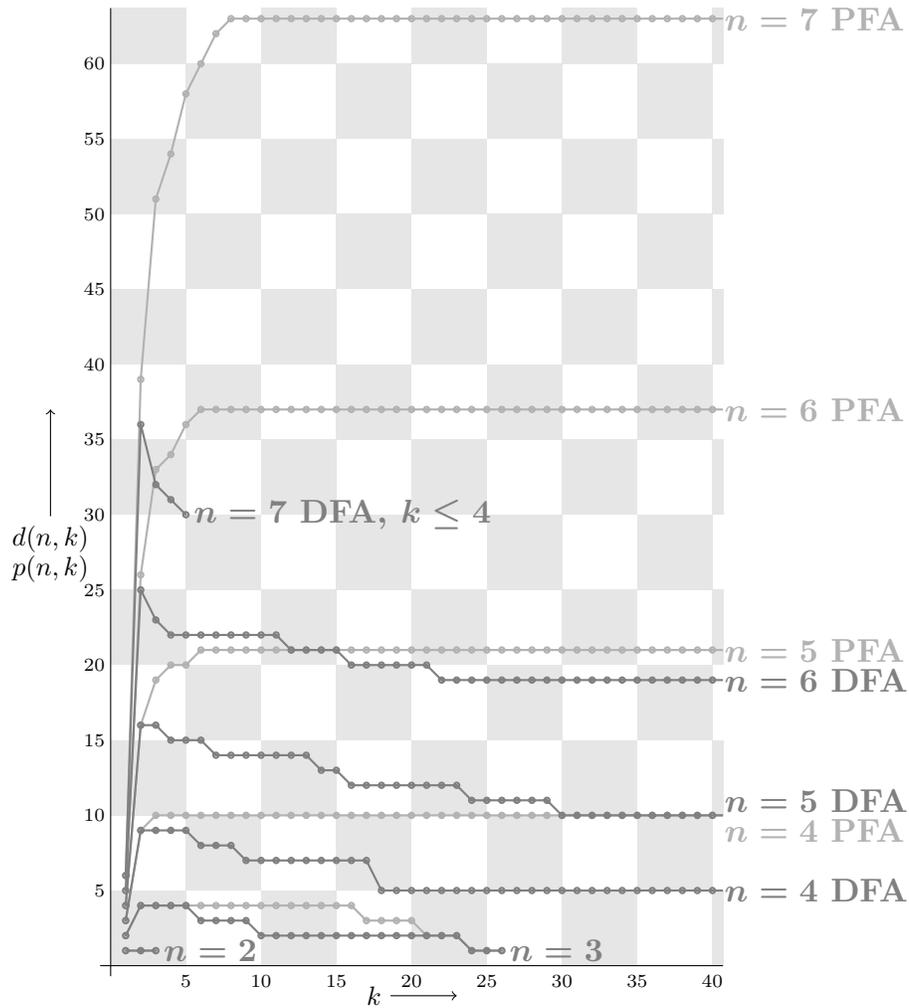
3.2. PFAs on at Most 7 States with Fixed Alphabet Size

Write $p(n, k)$ for the maximal shortest carefully synchronizing word length for a PFA on n states and k symbols. Computing the values of $p(n, k)$ for all $n \leq 7$ and all $k \leq 41$ is a lot more involved than computing $p(n)$ for all $n \leq 7$. We made several improvements to the algorithm to get it done, among which the following:

- (1) It appeared that most of the times where upper bound L'' needs to be determined, the PFA is already synchronizing. So we start with trying a breadth first search with bit vectors, and only compute L'' in the above-described way if the PFA is not synchronizing.
- (2) We estimate the number of required symbols after postprocessing (making symbols more complete) already before the postprocessing, and use this estimate to prune the search.
- (3) If the estimate on the number of required symbols is equal to the maximum allowed number of symbols, then for every extension \mathcal{B} of \mathcal{A} , the PFAs we get by postprocessing \mathcal{B} are contained in the PFAs we get by postprocessing \mathcal{A} directly. For that reason, we do not search further for extensions of \mathcal{A} in this case, but postprocess immediately. So the postprocessing is not only to reduce the number of symbols in this case, but also to obtain synchronization.

In the graph below, the values of $p(n, k)$ are plotted for all $n \leq 7$ and all $k \leq 40$ in light gray. Furthermore, the values of $d(n, k)$ for DFAs are plotted for all $n \leq 7$ and all $k \leq 40$ in dark gray, except the cases where $n = 7$ and $5 \leq k \leq 40$.

So, we see that for DFAs with $n \leq 7$ states, after having the maximum $d(n, k) = (n - 1)^2$ at $k = 2$, the values of $d(n, k)$ decrease for larger k . So it seems that for



this choice gives the largest possible synchronization lengths. In contrast, for binary PFAs the lengths grow slower than for general PFAs, although the growth is still exponential as we will see in Section 6.

Using breadth first search with bit vectors, combined with symmetry reduction on the states, we computed all possible carefully synchronization lengths of binary PFAs with $n \leq 9$ states. For the binary PFAs with $n = 10$ states, we additionally used multithreading and applied a few low level optimizations. One of the optimization techniques was to view the PFAs as CNFAs, namely by replacing undefined transitions by transitions to the whole set of states. The results are displayed below, where the maximum carefully synchronization lengths $p(n, 2)$ are in boldface. For comparison, we also added the known synchronization lengths for general PFAs.

n	binary DFA	proper binary PFA	PFA
2	1	1	1
3	1- 4	1-3	1-4
4	1- 9	1-7	1-10
5	1- 16	1-15	1-21
6	1-23, 25	1-23, 26	1-37
7	1-32, 36	1-33, 35-36, 39	1-63
8	1-44, 49	1-45, 48, 50, 52, 55	
9	1-52, 56-58, 64	1-63, 65, 68, 72- 73	
10	1-66, 72-74, 81	1-80, 82-84, 87, 89, 93- 94	

A notable feature in this table is that several gaps appear in the ranges of possible values. Unfortunately, we still lack a deeper understanding of this behaviour. For DFAs, existence of gaps has already been observed in [1,12,21] and has been studied further in [8].

4. Specific PFA Constructions

In this section we present two series of PFAs (parameterized by its size n) of special interest: they have quadratic shortest synchronizing word length exceeding $(n-1)^2$, for each n for which this is possible. Furthermore, they reach the maximum possible synchronization length for some low values of n . The constructed series fill up a void between the computations up to 7 or 10 states respectively, and the asymptotic results in the next sections.

Both series are closely related to Černý's DFAs. The first series is T_n on n states and three symbols; for this we give the full analysis which is quite straightforward. The second series is P_n on n states and two symbols. For this series the full analysis

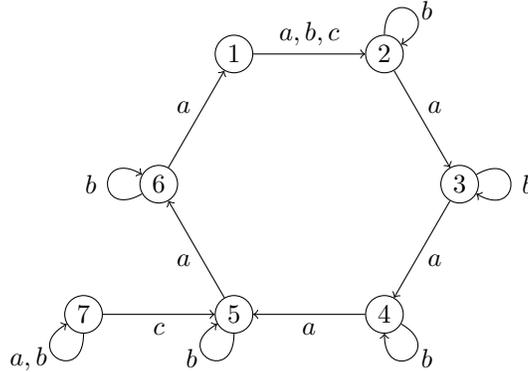
is much more involved; in this paper we give the construction and the results, but the full analysis leading to these results will be presented in a separate paper.

We start by T_n . For $n \geq 4$, T_n is defined to be the PFA on the n states $1, 2, \dots, n$ and the three symbols a, b, c such that

$$qa = \begin{cases} q+1 & 1 \leq q \leq n-2 \\ 1 & q = n-1 \\ n & q = n \end{cases} \quad qc = \begin{cases} 2 & q = 1 \\ \perp & 2 \leq q \leq n-1 \\ 2 + \lfloor \frac{n-1}{2} \rfloor & q = n \end{cases}$$

$$qb = \begin{cases} 2 & q = 1 \\ q & 2 \leq q \leq n \end{cases}$$

For $n = 3$ we take the same definition in which $nc = 2 + \lfloor \frac{n-1}{2} \rfloor$ is taken modulo $n-1$, so $3c = 1$. Note that for all n the PFA is obtained by extending C_{n-1} by an extra node n on which a and b act as the identity, and an extra symbol c that is only defined on 1 and n . The PFA T_n under consideration is depicted below for $n = 7$.



Theorem 2. *For every $n \geq 3$ the PFA T_n is carefully synchronizing with unique shortest synchronizing word $(ba^{n-2})^{n-2}cv$ of length $\frac{3(n-1)(n-2)}{2} + 1$, where $v = (a^{n-2}b)^{(n-2)/2}$ if n is even, and $v = a^{(n-3)/2}b(a^{n-2}b)^{(n-3)/2}$ if n is odd.*

Proof. First we show that the given word is carefully synchronizing. Write $Q = \{1, \dots, n\}$. Since C_{n-1} synchronizes with $(ba^{n-2})^{n-3}b$ ending in state 2, we obtain $Q(ba^{n-2})^{n-3}b = \{2, n\}$, followed by a^{n-2} yielding $Q(ba^{n-2})^{n-2} = \{1, n\}$, being the set on which c is defined, hence $Q(ba^{n-2})^{n-2}c = \{2, 2 + \lfloor \frac{n-1}{2} \rfloor\}$. It is easily checked that in C_{n-1} one has $\{2, 2 + \lfloor \frac{n-1}{2} \rfloor\}v = \{2\}$, passing all $\binom{n-1}{2}$ subsets of size 2 of $\{1, \dots, n-1\}$ exactly once, and decreasing the distance between the two elements by 1 every time a b from v is processed.

Conversely, let w be a shortest carefully synchronizing word for T_n . To include the state n in synchronization, w should contain a c , so write $w = w_1cw_2$ in which $w_1 \in \{a, b\}^*$. Since c should be defined on Qw_1 , we have $Qw_1 \subseteq \{1, n\}$. Ignoring state n , we obtain $\{1, \dots, n-1\}w_1 = \{1\}$ in C_{n-1} . Since the shortest prefix of

w_1 that synchronizes in C_{n-1} , synchronizes in state 2, $n - 2$ more a steps are needed to synchronize in state 1, so w_1 has length at least $n - 2$ plus the shortest synchronization length of C_{n-1} being $(n - 2)^2$, yielding $|w_1| \geq (n - 1)(n - 2)$. Note that the synchronizing word we gave satisfies $|w_1| = (n - 1)(n - 2)$. Since both 1 and n are contained in Qw_1 , we obtain $Qw_1c = \{2, 2 + \lfloor \frac{n-1}{2} \rfloor\}$.

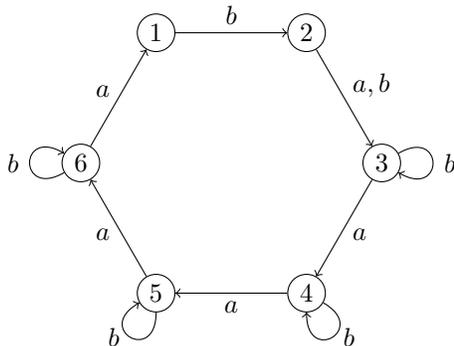
So until the singleton is obtained after applying w_2 to this set, all intermediate sets consist of two elements from $\{1, \dots, n - 1\}$. One checks that the distance between these two elements can only decrease by a b step, and only in the case the set contains state 1 and a state in $\{2, 3, \dots, \lfloor \frac{n-1}{2} \rfloor\}$. Synchronization is obtained if this distance becomes 0. Counting the numbers of b s and the numbers of intermediate a steps required to satisfy this requirement shows that v is the shortest candidate for w_2 . Hence no shorter carefully synchronizing word is possible than the one we gave. \square

Note that for $n = 3, 4$ the PFA T_n has the highest possible carefully synchronizing word length among all PFAs (4 and 10), while for $n = 5$ it is the highest possible among all PFAs on 3 symbols. Moreover, for all $n \geq 4$ it strictly exceeds $(n - 1)^2$. Furthermore, one can adapt symbol c of T_n , to obtain a PFA of which the synchronization length is any given number in $\{1, 2, \dots, \frac{3(n-1)(n-2)}{2}\}$.

A natural question is what are the worst cases for binary PFAs and small n . It turns out that we can find a similar class of binary PFAs, so with two symbols rather than three, and having similar properties. In the class of binary PFAs, Černý's example still is the worst possible for $n \leq 5$. For $6 \leq n \leq 10$, there is a unique binary PFA reaching the maximal length, being 26, 39, 55, 73 and 94 for $n = 6, 7, 8, 9, 10$ respectively. The first four of these PFAs are all members of a sequence P_n that we introduce now. Again it looks very much like Černý's sequence. For $n \geq 3$, P_n is defined by

$$qa = \begin{cases} \perp & q = 1 \\ q + 1 & 2 \leq q \leq n - 1 \\ 1 & q = n \end{cases} \quad qb = \begin{cases} q + 1 & 1 \leq q \leq 2 \\ q & 3 \leq q \leq n \end{cases}$$

The PFA P_n under consideration is depicted below for $n = 6$.



Although the construction of P_n is quite simple, the synchronization lengths show a somewhat curious pattern. Where sequences of DFAs in the literature generally give rise to quadratic or linear formula's, this is not the case for P_n . The lengths are quadratic in size, but no explicit quadratic formula for it exists. The synchronization length of P_n is strictly larger than $(n-1)^2$ for all $n \geq 6$. As b and ab act as C_{n-1} on $\{2, \dots, n\}$ it is easily seen that $b(b(ab)^{n-2})^{n-3}b$ is a carefully synchronizing word for P_n , but for $n \geq 5$ it is not a shortest one.

The synchronization length can be expressed in the Fibonacci numbers $\text{fib}(m)$ defined by $\text{fib}(0) = 0$, $\text{fib}(1) = 1$ and $\text{fib}(m) = \text{fib}(m-1) + \text{fib}(m-2)$ for $m \geq 2$. Let $\phi = \frac{1+\sqrt{5}}{2}$ be the golden ratio.

Theorem 3. *For $n \geq 3$, let m be the unique integer for which $\text{fib}(m-1) < n-2 \leq \text{fib}(m)$. If w is a shortest synchronizing word for P_n , then*

$$|w| = n^2 + mn - 5n - \text{fib}(m+1) - 2m + 8 = n^2 + \frac{n \log(n)}{\log(\phi)} + \Theta(n).$$

Furthermore, $|w| > (n-1)^2$ for $n \geq 6$.

The proof of this theorem is expected to appear in a forthcoming paper by Stijn Cambie and the first two authors. Below is a table of $|w|$ for small n .

n	3	4	5	6	7	8	9	10	11
$ w $	2	7	15	26	39	55	73	93	116

5. Exponential Bounds for PFAs on Three Symbols

In this section, we demonstrate our techniques to construct PFAs with only three symbols and exponential shortest synchronizing word length. These constructions are based on string rewrite systems. In the next section we will show a reduction to two symbols and the last section is devoted to more elaborate constructions that lead to sharper asymptotic results.

For any $k \geq 3$, we build a transitive PFA on $n = 3k$ states and three symbols, which is carefully synchronizing, and the shortest carefully synchronizing word has length $\Omega(\phi^{n/3})$, where $\phi = \frac{1+\sqrt{5}}{2} \approx 1.618$. The set of states is $Q = \{A_i, B_i, C_i \mid i = 1, \dots, k\}$. If a set $S \subseteq Q$ contains exactly one element of $\{A_i, B_i, C_i\}$ for every i , it can be represented by a string over $\{A, B, C\}$ of length k . The idea of our construction is that the PFA will mimic rewriting the string C^2A^{k-2} to the string $C^2A^{k-3}B$ with respect to the rewrite system R , which consists of the following three rules

$$BBA \rightarrow AAB, CBA \rightarrow CAB, CCA \rightarrow CCB.$$

The key argument is that this rewriting is possible, but requires an exponential number of steps. This is elaborated in the following lemma, in which we use \rightarrow_R for rewriting with respect to R , that is, $u \rightarrow_R v$, if and only if $u = u_1\ell u_2$ and $v = u_1ru_2$,

for strings u_1, u_2 and a rule $\ell \rightarrow r$ in R . Its transitive closure is denoted by \rightarrow_R^+ . Just as in the previous section, we write fib for the standard Fibonacci function. It is well-known that $\text{fib}(n) = \Theta(\phi^n)$.

Lemma 4. *For $k \geq 3$, we have $CCA^{k-2} \rightarrow_R^+ CCA^{k-3}B$. Furthermore, the smallest possible number of steps for rewriting CCA^{k-2} to a string ending in B , is exactly $\text{fib}(k) - 1$.*

Proof. For the first claim we do induction on k . For $k = 3$, we have $CCA \rightarrow_R CCB$. For $k = 4$, we have $CCAA \rightarrow_R CCBA \rightarrow_R CCAB$. For $k > 4$, applying the induction hypothesis twice, we obtain

$$CCA^{k-2} \rightarrow_R^+ CCA^{k-4}BA \rightarrow_R^+ CCA^{k-5}BBA \rightarrow_R CCA^{k-3}B.$$

For the second claim, we define the *weight* $W(u)$ of a string $u = u_1u_2 \cdots u_k$ over $\{A, B, C\}$ of length k by

$$W(u) = \sum_{i:u_i=B} (\text{fib}(i) - 1).$$

So every B on position i in u contributes $\text{fib}(i) - 1$ to the weight, and the other symbols have no weight.

Now we claim that $W(v) = W(u) + 1$ for all strings u, v with $u \rightarrow_R v$ and u, v only having C 's in the first two positions. Since the C 's only occur at positions 1 and 2, by applying $CCA \rightarrow CCB$, the weight increases by $\text{fib}(3) - 1 = 1$ by the creation of B on position 3, and by applying $CBA \rightarrow CAB$, it increases by $\text{fib}(4) - 1 - (\text{fib}(3) - 1) = 1$ since B on position 3 is replaced by B on position 4. By applying $BBA \rightarrow AAB$, the contributions to the weight $\text{fib}(i) - 1$ and $\text{fib}(i+1) - 1$ of the two B 's are replaced by $\text{fib}(i+2) - 1$ of the new B , which is an increase by 1 according to the definition of fib .

So this weight increases by exactly 1 at every rewrite step, hence it requires exactly $\text{fib}(k) - 1$ steps, to go from the initial string CCA^{k-2} of weight 0 to the weight $\text{fib}(k) - 1$ of a B symbol on the last position k , if that is the only B , and more steps if there are more B 's. \square

Now we are ready to define the PFA on $Q = \{A_i, B_i, C_i \mid i = 1, \dots, k\}$ and three symbols. The three symbols are a start symbol s , a rewrite symbol r and a cyclic

shift symbol c . The transitions are defined as follows (writing \perp for undefined):

$A_i s = B_i s = C_i s = C_i,$			for $i = 1, 2,$
$A_i s = B_i s = C_i s = A_i,$			for $i = 3, \dots, k,$
$A_1 r = \perp,$	$B_1 r = A_1,$	$C_1 r = C_1,$	
$A_2 r = \perp,$	$B_2 r = A_2,$	$C_2 r = C_2,$	
$A_3 r = B_3,$	$B_3 r = \perp,$	$C_3 r = B_2,$	
$A_i r = A_i,$	$B_i r = B_i,$	$C_i r = C_i,$	for $i = 4, \dots, k,$
$A_i c = A_{i+1},$	$B_i c = B_{i+1},$	$C_i c = C_{i+1},$	for $i = 1, \dots, k-1,$
$A_k c = A_1,$	$B_k c = B_1,$	$C_k c = C_1.$	

A shortest carefully synchronizing word starts by s , since r is not defined on all states and c permutes all states. After s , the set of reached states is $S(CCA^{k-2}) = \{C_1, C_2, A_3, \dots, A_k\}$. Here, for a string $u = a_1 a_2 \dots a_k$ of length k over $\{A, B, C\}$, we write $S(u)$ for the set of k states, containing A_i if and only if $a_i = A$, containing B_i if and only if $a_i = B$, and containing C_i if and only if $a_i = C$, for $i = 1, 2, \dots, k$. Note that for $x \in \{A, B, C\}$ and $v \in \{A, B, C\}^{k-1}$, we have $S(vx)c = S(xv)$, so c performs a cyclic shift on strings of length k .

The next lemma states that the symbol r indeed mimicks rewriting: applied on sets of the shape $S(u)$, up to cyclic shift it acts as rewriting on u with respect to R defined above.

Lemma 5. *Let u be a string of the shape CCw , where $w \in \{A, B\}^{k-2}$. If $u \rightarrow_R v$ for a string v , then $S(u)c^i r c^{k-i} = S(v)$ for some $i < k$.*

Conversely, if u does not end in B and there exists an i such that r is defined on $S(u)c^i$, then $u \rightarrow_R v$ for a string v of the shape CCw , where $w \in \{A, B\}^{k-2}$.

Proof. First assume that $u \rightarrow_R v$. If $u = u_1 BBAu_2$ and $v = u_1 AABu_2$, then let $i = |u_2| + 3$, so

$$\begin{aligned} S(u)c^i r c^{k-i} &= S(u_1 BBAu_2)c^i r c^{k-i} = S(BBAu_2 u_1) r c^{k-i} \\ &= S(AABu_2 u_1) c^{k-i} = S(u_1 AABu_2) = S(v). \end{aligned}$$

If $u = u_1 CBAu_2$ and $v = u_1 CABu_2$, then again let $i = |u_2| + 3$, so

$$\begin{aligned} S(u)c^i r c^{k-i} &= S(u_1 CBAu_2)c^i r c^{k-i} = S(CBAu_2 u_1) r c^{k-i} \\ &= S(CABu_2 u_1) c^{k-i} = S(u_1 CABu_2) = S(v). \end{aligned}$$

Finally, if $u = u_1 CCAu_2$ and $v = u_1 CCBu_2$, then $u_1 = \lambda$ and the result follows for $i = 0$.

Conversely, suppose that $S(u)c^i r$ is defined. Since $S(u)c^k = S(u)$, we may assume that $i < k$ and can write $u = u_1 u_2$, such that $|u_2| = i$. Then $S(u)c^i = S(w)$, where $w = u_2 u_1$. Write $w = a_1 a_2 \dots a_k$. Since $S(u_2 u_1) r$ is defined, we get $a_1 \neq A$, $a_2 \neq A$ and $a_3 \neq B$. Moreover, $a_1 = a_2 = a_3 = C$ does not occur since u only

contains 2 Cs, and $a_1a_2 = BC$ or $a_2a_3 = BC$ does not occur since u does not end in B . The remaining 3 cases are

$$a_1a_2a_3 = BBA, \quad a_1a_2a_3 = CBA, \quad \text{and} \quad a_1a_2a_3 = CCA,$$

where $a_1a_2a_3$ is replaced by the corresponding right hand side of the rule by the action of r . Then in $S(u)c^i r c^{k-i}$, the two Cs are on positions 1 and 2 again, and we obtain $S(u)c^i r c^{k-i} = S(v)$ for a string v of the given shape, satisfying $u \rightarrow_R v$ \square

Combining Lemmas 4 and 5 and the fact that $\text{fib}(n) = \Omega(\phi^n)$, we obtain the following.

Corollary 6. *There is a word w such that $S(CCA^{k-2})w = S(CCA^{k-3}B)$; the shortest word w for which $S(CCA^{k-2})w$ is of the shape $S(u)c^i$ for u ending in B has length $\Omega(\phi^k)$.*

Now we are ready to prove the lower bound:

Lemma 7. *If w is carefully synchronizing, then $|w| = \Omega(\phi^k)$.*

Proof. Assume that w is a shortest carefully synchronizing word. Then we already observed that the first symbol of w is s , and w yields $S(CCA^{k-2})$ after the first step in the power automaton. By applying only c -steps and r -steps, according to Lemma 5, only sets of the shape $S(u)c^i$ for which $CCA^{k-2} \xrightarrow{+}_R u$ can be reached, until u ends in B . In this process, each r -step corresponds to a rewrite step. Applying the third symbol s does not make sense, since then we go back to $S(CCA^{k-2})$. According to Corollary 6, in the power automaton at least $\Omega(\phi^k)$ steps are required to reach a set which is not of the shape $S(u)c^i$. So for reaching a singleton, the total number of steps is at least $\Omega(\phi^k)$. \square

Note that for the reasoning until now, the definition of $C_3r = B_2$ did not play a role, and by s, r all states were replaced by states having the same index. But after the last symbol of u has become B , this $C_3r = B_2$ will be applied, leading to a subset in which no state of the group A_3, B_3, C_3 occurs any more. We could have chosen $C_3r = A_2$ or $C_3r = C_2$ as well: it is just that $C_3r = B_2$ makes r injective, like c . Now we arrive at the main result of this section. Optimizations leading to sharper bounds will be presented in Section 7.

Proposition 8. *There exists a sequence of transitive carefully synchronizing PFAs with three symbols, n states and shortest synchronizing word length $\Omega(\phi^{n/3})$.*

Proof. Let $n = 3k + m$ with $m \in \{0, 1, 2\}$. Take our PFA on $3k$ states and select m states with more than one ingoing arrow. Split each of them into two states, each inheriting some of the ingoing arrows. This affects the injectivity of r and c , but the PFA remains transitive, and the bound for $3k$ states is maintained. The bound was proved in Lemma 7; it remains to prove that the PFA with $3k$ states is synchronizing, that is, it is possible to end up in a singleton in the power automaton.

Let w be the word from Corollary 6. Since $S(CCA^{k-2})w = S(CCA^{k-3}B)$ and the number of c 's in w is divisible by k , we have $C_1w = C_1$, $C_2w = C_2$, $A_3w = A_3, \dots, A_{k-1}w = A_{k-1}$, $A_kw = B_k$. Hence

$$\begin{aligned} \{A_1, B_1, C_1\}swcr &= \{C_1\}cr = \{C_2\} \subseteq \{A_1, B_1, C_1\}c, \\ \{A_2, B_2, C_2\}swcr &= \{C_2\}cr = \{B_2\} \subseteq \{A_2, B_2, C_2\}, \\ \{A_i, B_i, C_i\}swcr &= \{A_i\}cr = \{A_{i+1}\} \subseteq \{A_i, B_i, C_i\}c, \quad \text{for } i = 3, 4, \dots, k-1, \\ \{A_k, B_k, C_k\}swcr &= \{B_k\}cr = \{A_1\} \subseteq \{A_k, B_k, C_k\}c. \end{aligned}$$

So for all $i \neq 2$, $\{A_i, B_i, C_i\}swcr$ is contained in the cyclic successor $\{A_i, B_i, C_i\}c$ of $\{A_i, B_i, C_i\}$. $\{A_2, B_2, C_2\}swcr$ is just contained in $\{A_2, B_2, C_2\}$ itself. Since for any i , one can take the cyclic successor of $\{A_i, B_i, C_i\}$ at most $k-1$ times before ending up in $\{A_2, B_2, C_2\}$, we deduce that

$$\{A_i, B_i, C_i\}(swcr)^{k-1} \subseteq \{A_2, B_2, C_2\} \quad \text{for } i = 1, 2, \dots, k.$$

As $\{A_2, B_2, C_2\}s = \{C_2\}$, we obtain the carefully synchronizing word $(swcr)^{k-1}s$ of the PFA. \square

The word $(swcr)^{k-1}s$ is a lot longer than necessary. In fact, one can prove that only $O(k^2)$ c -steps and $O(k)$ r -steps and s -steps suffice after $swcr$.

6. Reduction to Two Symbols

In this section we construct PFAs with two symbols and exponential shortest carefully synchronizing word length. We do this by a general transformation to two-symbol PFAs, as was done before, e.g. in [23]. There a PFA on n states and m symbols was transformed to a PFA on mn states and two symbols, preserving synchronization length. In the next theorem, we improve this resulting number of states to $(m-1)n$ or even less, only needing a mild extra condition. Using this result, we reduce our 3-symbol PFA with synchronizing length $\Omega(\phi^{n/3})$ to a 2-symbol PFA with synchronizing length $\Omega(\phi^{n/5})$.

Theorem 9. *Let $P = (Q, \Sigma)$ be a carefully synchronizing PFA with $|Q| = n$, $|\Sigma| = m$, and shortest carefully synchronizing word length $f(n)$. Assume $s \in \Sigma$ and $Q' \subseteq Q$ satisfy the following properties.*

- (1) *there is some number p such that all symbols are defined on Qs^p for a complete symbol s ,*
- (2) *$qs = q$ for all $q \in Q'$, and*
- (3) *$qa = qb$ for all $q \in Q'$ and all $a, b \in \Sigma \setminus \{s\}$.*

Let $n' = n - |Q'|$. Then there exists a carefully synchronizing PFA on $n + n'(m-2)$ states and 2 symbols, with shortest carefully synchronizing word length at least $f(n)$. The new PFA is deterministic and/or transitive if P is.

Note that if $Q' = \emptyset$ then only requirement 1 remains, and the resulting number of states is $n + n'(m - 2) = (m - 1)n$.

Proof. Write $Q = \{1, 2, \dots, n\}$, $Q' = \{n' + 1, \dots, n\}$, and $\Sigma = \{s, a_1, \dots, a_{m-1}\}$. Let the states of the new PFA be $P_{1,j}$ for $j = 1, \dots, n$ and $P_{i,j}$ for $i = 2, \dots, m - 1$, $j = 1, \dots, n'$. Define the following two symbols a, b on these states:

$$P_{i,j}a = \begin{cases} P_{i+1,j}, & \text{if } i < m - 1, j \leq n', \\ P_{1,js}, & \text{if } i = m - 1, j \leq n', \\ P_{1,j}, & \text{if } i = 1, j > n'. \end{cases} \quad \begin{array}{cccc} P_{1,1} & \cdots & P_{1,n'} & P_{1,n'+1} \cdots P_{1,n} \\ P_{2,1} & \cdots & P_{2,n'} & \\ \vdots & & \vdots & \\ P_{m-1,1} & \cdots & P_{m-1,n'} & \end{array}$$

and $P_{i,j}b = P_{1,ja_i}$, for all $i = 1, \dots, m - 1$ and $j = 1, \dots, n$ for which $P_{i,j}$ exists and ja_i is defined.

If we arrange the states as indicated above, then on the leftmost n' columns, a moves the states one step downward if possible, and for the bottom row jumps to the top row and acts there as s . For the remainder of the top row a also acts as s (which is the identity). On the leftmost n' columns, the symbol b acts as a_i on row i and then jumps to the top line. For the remainder of the top row, all a_i act in the same way and b acts likewise.

Define $\psi(a_i) = a^{i-1}b$ for $i = 1, \dots, m - 1$, and $\psi(s) = a^{m-1}$. Then on the top line $\psi(a_i)$ acts in the same way as a_i in the original PFA. Similarly, $\psi(s)$ acts as s . On any other row, $\psi(s)$ acts as s , too. Since every symbol a_i is defined on qs^p for every $q \in Q$, we obtain that $\psi(s)^pb = a^{(m-1)p}b$ is defined on every state and ends up in the top row.

Assume that w is carefully synchronizing in the original PFA. Then by the above observations, $a^{(m-1)p}b\psi(w)$ is carefully synchronizing in the new PFA. Conversely, any carefully synchronizing word of the new PFA can be written as $\psi(w)a^j$, where $0 \leq j \leq m - 2$ and $\psi(w)$ is a concatenation of blocks of the form $\psi(l)$, $l \in \Sigma$. Now note that a^j can never synchronize two distinct states in the top row. Therefore, $\psi(w)$ synchronizes the top row and consequently w is synchronizing in the original PFA. Clearly $|\psi(w)a^j| \geq |w| \geq f(n)$. \square

We apply Theorem 9 to our basic construction with $3k$ states and $m = 3$ symbols; note that s, c are defined on all states and r is defined on Qs , so the requirements of Theorem 9 hold for $p = 1$. As r and c act differently on all states, the only option for Q' is $Q' = \emptyset$. Hence we obtain a carefully synchronizing PFA on $(m - 1)3k = 6k$ states and two symbols, with shortest carefully synchronizing word length $\Omega(\phi^k)$. For n being the number of states of the new PFA, this is $\Omega(\phi^{n/6})$.

However, instead of our three symbols s, c, r we also get careful synchronization on the three symbols s, c, rc with careful synchronization length of the same order. But then for $i = 4, \dots, k$ we have $A_i s = A_i$ and $A_i c = A_i rc$, so we may choose $Q' = \{A_4, \dots, A_k\}$ in Theorem 9, by which $n' = 3k - (k - 3) = 2k + 3$, yielding a

PFA on two symbols and $5k + 3$ states. This results in the following proposition, where for n not of the shape $5k + 3$ we remove up to four states from Q' .

Proposition 10. *There exists a sequence of transitive carefully synchronizing PFAs with two symbols, n states and shortest synchronizing word length $\Omega(\phi^{n/5})$.*

This result will be sharpened in the next section as well.

7. Main Asymptotic Results

In this section we discuss some further optimizations. First we extend the number of rewrite rules and then the number of letters in the system. These rewrite systems will be used to construct PFAs on two and three symbols for which we will derive asymptotic lower bounds for the synchronization length.

7.1. More Rewrite Rules

For any $h \geq 2$ we define a rewrite system R_h by taking $h + 1$ rewrite rules

$$C^i B^{h-i} A \rightarrow C^i A^{h-i} B \quad (1)$$

for $i = 0, \dots, h$. Then it is possible to construct a PFA that mimicks rewriting of the string $C^h A^{k-h}$ to $C^h A^{k-h-1} B$ in the system R_h . For $h = 2$ this coincides with our construction in Section 5, but for $h > 2$, this gives a better bound. The following lemma gives the number of steps needed. Note that $f_2(i)$ is equal to $\text{fib}(i) - 1$.

Lemma 11. *For $k \geq h + 1$, we have $C^h A^{k-h} \xrightarrow{+}_{R_h} C^h A^{k-h-1} B$. Furthermore, the smallest possible number of steps for rewriting $C^h A^{k-h}$ in the system R_h to a string ending in B is exactly $f_h(k)$, where $f_h(k)$ satisfies the recursion*

$$f_h(k) = \begin{cases} 0 & 1 \leq k \leq h \\ 1 + \sum_{j=1}^h f_h(k-j) & k \geq h + 1 \end{cases}$$

Proof. The proof is essentially analogous to the proof of Lemma 5. We define the weight $W(u)$ of a string $u = u_1 u_2 \dots u_k$ over $\{A, B, C\}$ by assigning weight w_i to a B on position i :

$$W(u) = \sum_{i: u_i=B} w_i.$$

Other symbols have zero weight. Now we want to choose w_i in such a way that every rewrite step increases the weight of a string by 1. This gives a recursion for w_i : to create a B in position i , we need u_j to be equal to B or C for all $i - h \leq j \leq i - 1$. After that, one extra rewrite step is needed. We start having already C 's in positions $1, \dots, h$. Therefore w_i satisfies

$$w_i = \sum_{j=\max\{h+1, i-h\}}^{i-1} w_j,$$

which means that $w_k = f_h(k)$ as defined in the lemma. By construction, to reach a string ending in B , exactly $f_h(k)$ rewrite steps are needed. \square

7.2. More Rewrite Symbols

Instead of just having A and B and rewriting the final A in a string into a B , we could take m symbols $A^{(1)}, \dots, A^{(m)}$. For convenience we will sometimes denote $A^{(1)}$ by A and $A^{(m)}$ by B . We take $(h+1)(m-1)$ rewrite rules

$$C^i B^{h-i} A^{(t)} \rightarrow C^i A^{h-i} A^{(t+1)}, \quad (2)$$

for $i = 0, \dots, h$ and $t = 1, \dots, m-1$. In this rewrite system $R_{h,m}$ the goal is to rewrite the string $C^h A^{k-h}$ into a string ending in B .

Lemma 12. *For $k \geq h+1$, we have $C^h A^{k-h} \xrightarrow{+}_{R_{h,m}} C^h A^{k-h-1} B$. Furthermore, the smallest possible number of steps for rewriting $C^h A^{k-h}$ in the system $R_{h,m}$ to a string ending in B is exactly $f_{h,m}(k)$, where $f_{h,m}(k)$ satisfies the recursion*

$$f_{h,m}(k) = \begin{cases} 0 & 1 \leq k \leq h \\ (m-1) \cdot \left(1 + \sum_{j=1}^h f_{h,m}(k-j)\right) & k \geq h+1 \end{cases}$$

Proof. We define the weight $W(u)$ of a string $u_1 u_2 \dots u_k$ by assigning weights to the symbols $A^{(t)}$ for $t \geq 2$:

$$W(u) = \sum_{t=2}^m \sum_{i: u_i = A^{(t)}} w_{i,t},$$

where $w_{i,t}$ is the weight of $A^{(t)}$ on position i . The symbols C and $A^{(1)} = A$ have zero weight. Again weights will be chosen such that every rewrite step increases the weight of a string by 1. Before we can replace a symbol $A^{(t)}$ in position i by $A^{(t+1)}$, we need u_j to be equal to $A^{(m)} = B$ or C for all $i-h \leq j \leq i-1$. After that, one extra rewrite step is needed. To replace a symbol $A^{(1)}$ in position i by $A^{(t)}$, this has to be repeated $t-1$ times. Therefore, we find the following recursion

$$w_{i,t} = (t-1) \left(1 + \sum_{j=\max\{h+1, i-h\}}^{i-1} w_{j,m} \right).$$

Then for all $k \geq h+1$, $W(C^h A^{k-h-1} B) = w_{k,m} = f_{h,m}(k)$. \square

7.3. Construction of the PFA on Three Symbols

Using the rewrite rules (2), we can construct a PFA $P_{h,m}^n$ on $n = (m+1)k$ states $A_i^{(1)}, \dots, A_i^{(m)}, C_i$ for $i = 1, \dots, k$. As before, we have a start symbol s , a rewrite symbol r and a cyclic shift symbol c . Let X denote any of the letters $A^{(1)}, \dots, A^{(m)}, C$. Then

$$X_i s = \begin{cases} C_i & i = 1, \dots, h \\ A_i^{(1)} & \text{otherwise} \end{cases} \quad X_i c = \begin{cases} X_1 & i = k \\ X_{i+1} & \text{otherwise} \end{cases} \quad (3)$$

Rewriting takes place in the states with indices $1 \leq i \leq h + 1$. We define it on $(m + 1)$ -tuples with index i by

$$(A_i^{(1)}r, \dots, A_i^{(m)}r, C_i r) = \begin{cases} (\perp, \dots, \perp, A_i^{(1)}, C_i) & i = 1, \dots, h \\ (A_i^{(2)}, \dots, A_i^{(m)}, \perp, A_{i-1}^{(m)}) & i = h + 1 \\ (A_i^{(1)}, \dots, A_i^{(m)}, C_i) & \text{otherwise.} \end{cases}$$

Lemma 13. *The PFA $P_{h,m}^n$ is carefully synchronizing and the shortest synchronizing word has length at least $f_{h,m}(n/(m + 1))$.*

Proof. Let Q be the state set of the PFA $P_{h,m}^n$. For a string $u = u_1 \dots u_k$ over $\{A^{(1)}, \dots, A^{(m)}, C\}$, we define $S(u) \subseteq Q$ in such a way that $X_i \in S(u)$ if and only if $u_i = X$. Then $Qs = S(C^h A^{k-h})$. Every application of the symbol r to a set $S(u)$ corresponds to application of a rewrite rule to u . As long as the string does not end in B , no other changes are possible, except for cyclic shifts and resetting to $C^h A^{k-h}$. To reach the set $S(C^h A^{k-h-1}B)$, we need at least a word of length $f_{h,m}(k) = f_{h,m}(n/(m + 1))$.

To see that $P_{h,m}^n$ is synchronizing, let w be such that $Qsw = S(C^h A^{k-h-1}B)$, and let

$$Q_i = \{A_i^{(1)}, \dots, A_i^{(m)}, C_i\}.$$

Then $Q_i swcr \subseteq Q_{(i+1) \bmod k}$ for all $i \neq h$, and $Q_h swcr \subseteq Q_h$. Consequently, $Q(swcr)^{k-1} s = \{C_h\}$ so that $(swcr)^{k-1} s$ is synchronizing. \square

7.4. Asymptotic Lower Bound for PFAs on Three Symbols

Theorem 14. *There exists a sequence of transitive carefully synchronizing PFAs with three symbols, n states and shortest carefully synchronizing word length*

$$\Omega\left(2^{\frac{2}{5}n - \log_2(n)}\right) = \Omega\left(\frac{2^{2n/5}}{n}\right).$$

Proof. As before, we can reduce to the case where $m + 1 \mid n$. For this case, we analyze the recursion of Lemma 12 and choose $h \geq 2$ and $m \geq 2$ dependent on n in such a way that $f_{h,m}(n/(m + 1))$ is maximal. First note that the recursive equations can be rewritten to a homogeneous system, by taking $g_{h,m}(k) = f_{h,m}(k) + \frac{m-1}{(m-1)h-1}$:

$$g_{h,m}(k) = \begin{cases} \frac{m-1}{(m-1)h-1} & k = 1, \dots, h \\ (m-1) \sum_{j=1}^h g_{h,m}(k-j) & k \geq h+1 \end{cases} \quad (4)$$

The case $m = h = 2$ gives Fibonacci's sequence. The general homogeneous recurrence relation has characteristic equation

$$x^h = (m-1)(x^{h-1} + x^{h-2} + \dots + 1) = (m-1) \cdot \frac{x^h - 1}{x - 1},$$

provided $x \neq 1$. It can be rewritten as $m - x = (m - 1)x^{-h}$, having a solution close to m . Indeed, for given $\varepsilon > 0$, we can choose h large enough so that there is a solution $\phi_{h,m}$ satisfying $m - \varepsilon \leq \phi_{h,m} < m$. This gives exponential growth with rate at least $m - \varepsilon$, which we will prove by induction. For $k = h + 1, \dots, 2h$, we have

$$g_{h,m}(k) \geq f_{h,m}(k) = (m - 1)m^{k-h-1} \geq \frac{m - 1}{(m - \varepsilon)^{h+1}} \cdot (m - \varepsilon)^k. \quad (5)$$

For $k > 2h$, assuming the above inequality for $k' < k$, we obtain

$$\begin{aligned} g_{h,m}(k) &= (m - 1) \sum_{j=1}^h g_{h,m}(k - j) \\ &\geq \frac{(m - 1)^2}{(m - \varepsilon)^{h+1}} \cdot (m - \varepsilon)^{k-1} \cdot \sum_{j=1}^h (m - \varepsilon)^{1-j} \\ &\geq \frac{(m - 1)^2}{(m - \varepsilon)^{h+1}} \cdot \frac{(m - \varepsilon)^k}{m - \varepsilon} \cdot \frac{1 - (m - \varepsilon)^{-h}}{1 - (m - \varepsilon)^{-1}} \\ &= \frac{m - 1}{(m - \varepsilon)^{h+1}} \cdot (m - \varepsilon)^k \cdot \frac{(m - 1) - (m - 1)(m - \varepsilon)^{-h}}{(m - 1) - \varepsilon}. \end{aligned}$$

In order to prove (5) for all $k > h$, the second fraction on the right hand side must be at least 1, which is equivalent to

$$(m - 1)(m - \varepsilon)^{-h} \leq \varepsilon.$$

So we take

$$h = \left\lceil \log_{m-\varepsilon} \left(\frac{m-1}{\varepsilon} \right) \right\rceil.$$

This implies $(m - \varepsilon)^{h-1} < (m - 1)/\varepsilon$, which we substitute in (5) to get for $k > h$

$$g_{h,m}(k) \geq \frac{m - 1}{(m - \varepsilon)^2} \cdot \frac{\varepsilon}{m - 1} \cdot (m - \varepsilon)^k \geq \frac{1}{m^2} \cdot \varepsilon \cdot (m - \varepsilon)^k. \quad (6)$$

The PFA that has to be constructed has $n = (m + 1)k$ states, so to find the growth in n , we substitute $k = \frac{n}{m+1}$ in (6). The following indirect argument shows that this is a valid choice, i.e. that $k > h$. If $k \leq h$, then the right hand side of (6) would be less than 1. Since our choice of k will lead to a lower bound greater than 1, we deduce that $k > h$. The best choice for m is $m = 4$, since this maximizes the growth rate $m^{1/(m+1)}$. This choice gives

$$g_{h,m} \left(\frac{n}{m+1} \right) \geq \frac{1}{16} \cdot \varepsilon \cdot (4 - \varepsilon)^{n/5}.$$

Finally, we want to choose ε . The right hand side has a maximum at $\varepsilon = 20/(n + 5)$. Note that this means that we rewrite a string of length k by substituting blocks of length $h + 1$ proportional to $\log(k)$. This choice leads to

$$g_{h,m} \left(\frac{n}{m+1} \right) \geq \frac{20}{16(n+5)} \left(4 - \frac{20}{n+5} \right)^{n/5} \geq \frac{5}{4(n+5)} \cdot 4^{n/5} \cdot \left(1 - \frac{5}{n} \right)^{n/5},$$

in which the last factor is bounded from below by a positive constant. Therefore

$$g_{h,m} \left(\frac{n}{m+1} \right) = \Omega \left(\frac{4^{n/5}}{n} \right) = \Omega \left(2^{\frac{2}{5}n - \log_2(n)} \right).$$

Since $f_{h,m}(n/(m+1))$ has the same growth rate, Lemma 13 gives the result. \square

7.5. Construction of the Binary PFA

To obtain an asymptotic lower bound for binary PFAs, we will use the reduction technique from Section 6. Before doing so, we will slightly tune the construction of $P_{h,m}^n$ so that we obtain a bigger set Q' in Theorem 9. Let $\tilde{P}_{h,m}^n$ be the PFA with symbols c, rc and a modified start symbol s' , defined by

$$\begin{aligned} X_i s' &= C_i & i &= 1, \dots, h \\ A_i^{(t)} s' &= A_i^{(t)} \text{ and } C_i s' = A_i^{(1)} & i &= h+2, \dots, k-4; t = 1, \dots, m \\ X_i s' &= A_i^{(1)} & i &= h+1 \text{ and } i = k-3, \dots, k, \end{aligned}$$

where X stands for any of the symbols $C, A^{(1)}, \dots, A^{(m)}$.

Lemma 15. *The automaton $\tilde{P}_{h,m}^n$ is carefully synchronizing and its shortest synchronizing word has length at least $\Omega \left(2^{\frac{2}{5}n - \log_2(n)} \right)$.*

Proof. Since $s = (cs')^k$, we deduce that $\tilde{P}_{h,m}^n$ is synchronizing. Now the set Qs' corresponds to a collection of strings of length k . More precisely, all strings of the form $u = C^h A u_{h+2} \dots u_{k-4} A^4$ with $u_j \in \{A^{(1)}, \dots, A^{(m)}\}$ for $j = h+2, \dots, k-4$. In this collection, the string with maximal weight is $u_{\max} := C^h A B^{k-h-5} A^4$. The number of steps to synchronize $\tilde{P}_{h,m}^n$ is at least the number of steps needed to rewrite u_{\max} into a string ending in B . To show that this is of the same order as rewriting a string of weight zero, it suffices to show that

$$W(u_{\max}) \leq \frac{1}{2} W(C^h A^{k-h-1} B). \quad (7)$$

First we prove that

$$W(C^h B^i A^{k-h-i}) \leq W(C^h A^{i+1} B A^{k-h-i-2}), \quad 0 \leq i \leq k-h-2. \quad (8)$$

For $i = 0$ and $i = 1$ this is clear by construction of the weights. By induction it then follows that

$$\begin{aligned} W(C^h B^{i+2} A^{k-h-i-2}) &= W(C^h B^i A^{k-h-i}) + W(C^h A^i B^2 A^{k-h-i-2}) \\ &\leq W(C^h A^{i+1} B A^{k-h-i-2}) + W(C^h A^{i+2} B A^{k-h-i-3}) \\ &= W(C^h A^{i+1} B^2 A^{k-h-i-3}) \\ &\leq W(C^h A^{i+3} B A^{k-h-i-4}), \end{aligned}$$

so (8) holds if $0 \leq i \leq k-h-2$. Taking $i = k-h-4$, we obtain

$$W(u_{\max}) \leq W(C^h B^{k-h-4} A^4) \leq W(C^h A^{k-h-3} B A^2) \leq \frac{1}{2} W(C^h A^{k-h-1} B),$$

proving (7), so that we get a lower bound of the same order as for $P_{h,m}^n$. \square

7.6. Asymptotic Lower Bound for Binary PFAs

We will apply Theorem 9 to the sequence $\tilde{P}_{h,m}^n$ to derive an asymptotic lower bound for binary PFAs.

Theorem 16. *There exists a sequence of carefully synchronizing PFAs with two symbols, n states and shortest carefully synchronizing word length*

$$\Omega\left(2^{\frac{1}{3}n - \frac{3}{2}\log_2(n)}\right) = \Omega\left(\frac{2^{n/3}}{n\sqrt{n}}\right).$$

Proof. Consider the PFA $\tilde{P}_{h,m}^n$. We check the conditions for Theorem 9. The symbol s' is complete and all symbols are defined on Qs' . Define the set of states Q' by

$$Q' = \{C_1, \dots, C_h\} \cup \{A_i^{(t)} \mid h+2 \leq i \leq k-4, 1 \leq t \leq m\},$$

fulfilling all conditions of Theorem 9. The reduction in this case gives a binary PFA on $N = 2n - |Q'| = 2(m+1)k - h - m(k-h-5) = (m+2)k + (m-1)h + 5m$ states so that

$$k = \frac{N - (m-1)h - 5m}{m+2}.$$

For $g_{h,m}(k)$ we still have the lower bound $\varepsilon \cdot (m-\varepsilon)^k/m^2$ as in (6). This time the main order term is $m^{N/(m+2)}$, which is again maximized by taking $m=4$. For ε , the best choice is $\varepsilon = 24/(N+6)$, which means $h = \log_{m-\varepsilon}(N) + O(1)$. Finally, we conclude that the length of the shortest synchronizing word for $\tilde{P}_{h,m}^n$ is bounded by

$$\begin{aligned} \Omega(\varepsilon \cdot (m-\varepsilon)^k) &= \Omega\left(\varepsilon \cdot (4-\varepsilon)^{N/6} \cdot (4-\varepsilon)^{-\log_{4-\varepsilon}(N)/2}\right) \\ &= \Omega\left(\frac{4^{N/6}}{N\sqrt{N}}\right) = \Omega\left(2^{\frac{N}{3} - \frac{3}{2}\log_2(N)}\right). \end{aligned}$$

If the number of states is not of the form $(m+2)k + (m-1)h + 5m = 6k + 3h + 20$, then we remove some states from Q' , just as before. \square

7.7. PFAs with a Single Undefined Transition

The PFA $P_{h,m}^n$ as defined in Section 7.3 has $h(m-1) + 1$ undefined transitions. In this section we present a variation on the theme, showing that a single undefined transition suffices to get exponential synchronizing word lengths. First note that the recursion in Lemma 12 gives exponential growth for $h=1$, provided $m \geq 3$. In this case, the recursion reduces to $f_{1,m}(k) = (m-1)(1 + f_{1,m}(k-1))$ with $f_{1,m}(1) = 0$. By a straightforward inductive argument, it follows that

$$f_{1,m}(k) = \frac{m-1}{m-2} \left((m-1)^k - 1 \right).$$

We will use the system (2) to rewrite $C(A^{(1)})^{k-1}$ into a string ending in $A^{(m)}$. We extend the rewrite system for $h = 1$ and $m \geq 3$ with $m - 2$ letters $A^{(m+1)}, \dots, A^{(2m-2)}$. We also extend the set of rewrite rules to

$$A^{(s)}A^{(t)} \rightarrow A^{(s-(m-1))}A^{(t+1)} \quad \text{and} \quad CA^{(t)} \rightarrow CA^{(t+1)}$$

for $t = 1, \dots, 2m - 3$ and $s \geq m$. Furthermore, we close the system cyclically by

$$A^{(s)}A^{(2m-2)} \rightarrow A^{(s-(m-1))}A^{(1)} \quad \text{and} \quad CA^{(2m-2)} \rightarrow CA^{(1)}.$$

for $s \geq m$. Just as before, for any $t \geq 1$, the weight of $A^{(t)}$ on some position is $t - 1$ times the weight of $A^{(2)}$ on the same position, and we see that the new rewrite rules either decrease the weight of a string or increase it by at most 1. So these extensions will not reduce the number of rewrite steps needed.

Now we build a PFA on $n = (2m - 1)k$ states $A_i^{(0)}, A_i^{(1)}, \dots, A_i^{(2m-2)}$ for $i = 1, \dots, k$ mimicking this rewrite system. Just as before, we make the rewrite symbol r injective, but this time it makes the construction slightly more complicated than necessary. The idea will be that a letter $A^{(t)}$ on position i in the string corresponds to the set of states

$$S_i^{(t)} := \begin{cases} \{A_i^{(t)}, \dots, A_i^{(t+m-2)}\} & 1 \leq t \leq m, \\ \{A_i^{(1)}, \dots, A_i^{(t-m)}, A_i^{(t)}, \dots, A_i^{(2m-2)}\} & m+1 \leq t \leq 2m-2. \end{cases}$$

Furthermore, the letter C on position i will correspond to either

$$S_i^{(0)} := \{A_i^{(0)}, A_i^{(1)}, \dots, A_i^{(m-2)}\} \quad \text{or} \quad \bar{S}_i^{(0)} := \{A_i^{(0)}, A_i^{(m)}, \dots, A_i^{(2m-3)}\}.$$

We define the start symbol s and the cyclic shift symbol c by

$$A_i^{(t)}s = \begin{cases} A_i^{(t \bmod (m-1))} & (m-1) \nmid t \text{ or } i = 1 \\ A_i^{(m-1)} & \text{otherwise} \end{cases} \quad A_i^{(t)}c = \begin{cases} A_1^{(t)} & i = k \\ A_{i+1}^{(t)} & \text{otherwise} \end{cases}$$

With this definition of the start symbol s , we have $Qs = S_1^{(0)} \cup \bigcup_{i=2}^k S_i^{(1)}$, representing the string $A^{(0)}(A^{(1)})^{k-1}$. The (injective) rewrite symbol is defined by

$$(A_i^{(0)}r, \dots, A_i^{(2m-2)}r) = \begin{cases} (A_1^{(0)}, A_1^{(m)}, \dots, A_1^{(2m-3)}, \perp, A_1^{(1)}, \dots, A_1^{(m-1)}) & i = 1 \\ (A_1^{(2m-2)}, A_2^{(2)}, \dots, A_2^{(2m-2)}, A_2^{(1)}) & i = 2 \\ (A_i^{(0)}, \dots, A_i^{(2m-2)}) & i \geq 3 \end{cases}$$

This implies that r acts on the sets $S_i^{(t)}$ for $i = 1, 2$ as

$$S_1^{(t)}r = \begin{cases} \bar{S}_1^{(0)} & t = 0 \\ \perp & 1 \leq t \leq m-1 \\ S_1^{(t-(m-1))} & m \leq t \leq 2m-2, \end{cases} \quad S_2^{(t)}r = \begin{cases} S_2^{(t+1)} & 1 \leq t \leq 2m-3 \\ S_2^{(1)} & t = 2m-2, \end{cases}$$

Since $\bar{S}_1^{(0)}r = S_1^{(0)}$ in addition, the states with indices $i = 1$ and $i = 2$ exactly mimic the rewrite rules. The action of r onto $S_2^{(0)}$ or $\bar{S}_2^{(0)}$ does not give a set of the form

$S_i^{(t)}$, but can only be applied after reaching a string ending in $A^{(t)}$ for some $t \geq m$. For $i \geq 3$, we have $S_i^{(t)}r = S_i^{(t)}$ for every t , and $\bar{S}_i^{(0)}r = \bar{S}_i^{(0)}$.

This construction leads to the following theorem.

Theorem 17. *There exists sequences of carefully synchronizing PFAs with only one undefined transition and shortest carefully synchronizing word length*

- $\Omega(3^{n/7})$ for PFAs on three symbols,
- $\Omega(2^{n/5})$ for PFAs on two symbols,

where n is the number of states.

Proof. First we argue that the PFA constructed above is synchronizing. Let w be a word to rewrite $C(A^{(1)})^{k-1}$ into a string ending in $A^{(m)}$. Write Q_i for the set of states with subindex i for each i . Then $Q_i c^{k-1} swc \subseteq Q_i$ for all i . Furthermore, r is defined on $Qc^{k-1}swc$, so let $v = c^{k-1}swcr$. Then $Q_i v \subseteq Q_i$ for all $i \neq 2$.

To investigate $Q_2 v$, we group states modulo $m-1$, by defining $B_i^{(0)} = \{A_i^{(0)}, A_i^{(m-1)}, A_i^{(2m-2)}\}$, and $B_i^{(t)} = \{A_i^{(t)}, A_i^{(t+m-1)}\}$ for all $t \neq 0$. Then $B_1^{(t)}sw \subseteq B_1^{(t)}$, so

$$B_2^{(t)}v \subseteq B_2^{(t)}r \subseteq B_2^{((t+1) \bmod (m-1))}$$

for all $t \neq 0$, and

$$B_2^{(0)}v \subseteq \{A_2^{(0)}\}r \subseteq B_1^{(0)}.$$

Consequently, $Q_2 v^{m-1} \subseteq Q_1$.

Furthermore, it follows by induction that $Q(v^{m-1}c)^{k-1} \subseteq Q_1$. From $B_k^{(0)}w \subseteq B_k^{(0)}$, we infer that $B_1^{(0)}v \subseteq B_1^{(0)}$. Hence $Q(v^{m-1}c)^k \subseteq B_1^{(0)}$. As $B_1^{(0)}s = \{A_1^{(0)}\}$, we conclude that $(v^{m-1}c)^k s$ is a synchronizing word. So we have a synchronizing n -state PFA with synchronizing word length

$$\Omega((m-1)^k) = \Omega((m-1)^{\frac{n}{2m-1}})$$

The best choice is $m = 4$, leading to the lower bound $\Omega(3^{n/7})$. If n is not of the form $7k$, then we can split up states just as before, but we must not split up state $A_1^{(m-1)}$.

For the binary construction with a single undefined transition, we proceed in the spirit of Section 7.5. We take symbols c and rc and use an adapted start symbol s' such that $A_i^{(t)}s' = A_i^{(t)}$ for $1 \leq t \leq 2m-2$ and $3 \leq i \leq k-4$. Now we can apply Theorem 9 with a set Q' of size $(2m-2)(k-6) + O(1)$. This gives a PFA on $n = 2mk$ states and a lower bound

$$\Omega((m-1)^k) = \Omega((m-1)^{\frac{n}{2m}}).$$

The best choice is $m = 5$, giving $\Omega(4^{n/10}) = \Omega(2^{n/5})$. \square

8. Conclusions

For every n , we constructed PFAs on n states and two or three symbols for which careful synchronization is forced to mimic rewriting with respect to a string rewrite system. These systems require an exponential number of steps to reach a string of a particular shape. The resulting exponential lengths are much larger than the cubic upper bound for synchronization of DFAs. We show that for $n = 4$ the shortest synchronization length for a PFA already can exceed the maximal shortest synchronization length for a DFA.

For $n \leq 7$ we found greatest possible shortest synchronization lengths, both for DFAs and PFAs, where for DFAs until now this was only fully investigated for $n \leq 4$, that is, by not assuming any bound on the number of symbols. For these n , we identify PFAs reaching the maximal length. These extreme cases require up to eight symbols, where for DFAs the maximal lengths are generally attained by binary examples.

Besides the proof of Theorem 3, several results which are related to those in this paper were not selected in this paper. One of those results is a generalization of the class P_n in Section 4. The other results have been gathered in [4].

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