

The Tractability Frontier for NFA Minimization*

Henrik Björklund and Wim Martens

TU Dortmund

henrik.bjoerklund@udo.edu

wim.martens@udo.edu

Abstract. We essentially show that minimizing finite automata is NP-hard as soon as one deviates from the class of deterministic finite automata. More specifically, we show that minimization is NP-hard for all finite automata classes that subsume the class that is unambiguous, allows at most one state q with a non-deterministic transition for at most one alphabet symbol a , and is allowed to visit state q at most once in a run. Furthermore, this result holds even for automata that only accept finite languages.

1 Introduction

The regular languages are immensely important, not only in theoretical computer science, but also in practical applications. When using regular languages in practice, the developer is often faced with a trade-off between the descriptive complexity and the complexity of optimization. Concretely, it has been known for a long time that there are regular languages for which non-deterministic finite automata (NFAs) can provide an exponentially more succinct description than deterministic finite automata (DFAs) [13]. On the other hand, many decision problems that are solvable in polynomial time for DFAs, i.e., equivalence, inclusion, and universality, are computationally hard for NFAs.

The choice of a representation mechanism can therefore be crucial. If the set of regular languages used in an application is relatively constant, membership tests are the main language operations, and economy of space is an issue, NFAs are probably the right choice. If, on the other hand, the languages change frequently, and inclusion or equivalence tests are frequent, DFAs may be more attractive.

Since both NFAs and DFAs have their disadvantages, a lot of effort has been spent on trying to find intermediate models, i.e., finite automata that have some *limited* form of non-determinism. A rather successful intermediate model is the class of *unambiguous* finite automata (UFAs).¹ While in general still being exponentially more succinct than an equivalent DFA for the same language, static analysis questions such as inclusion and equivalence can be solved in PTIME for UFAs [17]. However, UFAs do not allow for tractable state minimization [11].

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¹ An automaton is unambiguous if it has at most one accepting run for each word.

Therefore, the question whether there are good intermediate models between DFAs and NFAs needs to be revisited for state minimization. The main result of this paper is that, probably, no such good models exist. Even the tiniest bit of non-determinism makes the minimization problem NP-hard.

Every undergraduate computer science curriculum teaches its students how to minimize a DFA in polynomial time. In contrast, minimizing unrestricted NFAs is PSPACE-complete [18]. The minimization problem for automata with varying degrees of non-determinism was studied in a seminal paper by Jiang and Ravikumar in 1993 [11]. Among other results, they thoroughly investigated the minimization problem for UFAs. They showed the following.

- Given a UFA, finding the minimal equivalent UFA is NP-complete.
- Given a DFA, finding the minimal equivalent UFA is NP-complete.
- Given a DFA, finding the minimal equivalent NFA is PSPACE-complete.

Minimization problems have even been studied for automata with unary alphabets; see, e.g., [10, 5].

Recently, Malcher [12] improved on the results of Jiang and Ravikumar in the sense that he showed that finite automata with quite a small amount of non-determinism are hard to minimize. More precisely, he showed the following:

- (a) Minimization is NP-complete for automata that can non-deterministically choose between a fixed number of initial states, but are otherwise deterministic.
- (b) Minimization is NP-complete for non-deterministic automata with a constant number of computations for each string.²

Whereas Malcher made significant progress in showing that minimization is hard for non-deterministic automata, he was not yet able to answer the entire question. Therefore, he asks whether there are relaxations of the deterministic automata model *at all* for which minimization is tractable as an important open problem. In this respect, he mentions the class of automata with at most two computations for each string and the two classes (a) and (b) above with the added restriction of unambiguousness as important remaining cases. In the present paper, we settle these open questions and provide a uniform NP-hardness proof for all classes of automata mentioned above.

In brief, we define a class δ NFA of automata that are unambiguous, have at most two computations per string, and have at most one state q with two outgoing a -transitions, for at most one symbol a . Then, we show that minimization is NP-hard for all classes of finite automata that include δ NFAs, and show that these hardness results can also be adapted to the setting of unambiguous automata that can non-deterministically choose between two start states, but are deterministic everywhere else. This solves the open cases mentioned by Malcher. On the other hand, there *are* relaxations of the deterministic automaton model that allow tractable minimization. We show that, if we add to the definition of

² Actually, he showed this for automata with constant *branching*, which is slightly different from the number of computations; see Section 2.1.

δ NFAs that each word should have at most one *rejecting* computation, minimization becomes tractable again. However, the minimal automata in this class are the DFAs, so this class is not likely to be very useful in practice.

Other related work. An overview of state and transition complexity of NFAs can be found in [15]. Known results about the trade-off between amount of non-determinism and descriptive complexity are surveyed in [2]. The problems of producing small NFAs from regular expressions has been considered in [9, 16]. There has also been some work on approximating minimal NFAs [3, 4, 6], basically showing that approximation is very hard. The work of Hromkovic et al. [8] about measuring non-determinism in finite automata will be relevant to us in Section 2. Since the minimization problem for NFAs is hard, the bisimulation minimization problem has also been considered; see, e.g., [14, 1].

Due to space restrictions, some proofs and parts of proofs have been omitted, and will appear in the full version of the paper.

2 Preliminaries

By Σ we always denote a finite alphabet. A (*non-deterministic*) *finite automaton (NFA)* over Σ is a tuple $A = (\text{States}(A), \text{Alpha}(A), \text{Rules}(A), \text{init}(A), \text{Final}(A))$, where $\text{States}(A)$ is its finite set of states, $\text{Alpha}(A) = \Sigma$, $\text{init}(A) \in \text{States}(A)$ is its initial state, $\text{Final}(A) \subseteq \text{States}(A)$ is its set of final states, and $\text{Rules}(A)$ is a set of transition rules of the form $q_1 \xrightarrow{a} q_2$, where $q_1, q_2 \in \text{States}(A)$ and $a \in \Sigma$. The *size* of an automaton is $|\text{States}(A)|$, i.e., its number of states. A finite automaton is *deterministic* if, for each $q_1 \in \text{States}(A)$ and $a \in \text{Alpha}(A)$, there is at most one $q_2 \in \text{States}(A)$ such that $q_1 \xrightarrow{a} q_2 \in \text{Rules}(A)$. By DFA we denote the class of deterministic finite automata.

A *run*, or *computation*, r of A on a word $w = a_1 \cdots a_n \in \Sigma^*$ is a string $q_1 \cdots q_n \in \text{States}(A)^*$ such that $\text{init}(A) \xrightarrow{a_1} q_1 \in \text{Rules}(A)$ and, for each $i = 1, \dots, n-1$, $q_i \xrightarrow{a_{i+1}} q_{i+1} \in \text{Rules}(A)$. The run is *accepting* if $q_n \in \text{Final}(A)$. The *language of A* is the set of words w such that there exists an accepting run of A on w . A finite automaton A is *unambiguous* if, for each string w , there exists at most one accepting run of A on w .

Let \mathcal{N}_1 and \mathcal{N}_2 be two classes of NFAs. We say that $\mathcal{N}_1 \subseteq \mathcal{N}_2$ if each automaton in \mathcal{N}_1 also belongs to \mathcal{N}_2 . For example, $\text{DFA} \subseteq \text{NFA}$.

2.1 Notions of Non-Determinism

We recall some standard measures of non-determinism in a finite automaton. For a state q and an alphabet symbol a , the *degree of non-determinism* of a pair (q, a) , denoted by $\text{degree}(q, a)$ is the number k of different states q_1, \dots, q_k such that, for all $1 \leq i \leq k$, $q \xrightarrow{a} q_i \in \text{Rules}(A)$. We say that A has *degree of non-determinism k* , denoted by $\text{degree}(A) = k$, if $\text{degree}(q, a) \leq k$ for every $(q, a) \in \text{States}(A) \times \text{Alpha}(A)$, and there is at least one pair (q, a) such that $\text{degree}(q, a) = k$.

The *branching* of an automaton is intuitively defined as the maximum product of the degrees of non-determinism over states in a possible run. Formally, the branching of A on a word $w = a_1 \cdots a_n$ is $\text{branch}_A(w) = \max\{\prod_{i=1}^n \text{degree}(q_{i-1}, a_i) \mid q_1 \cdots q_n \text{ is a run of } A \text{ on } a_1 \cdots a_n, \text{ and } \text{init}(A) = q_0\}$. The *branching of A* , denoted $\text{branch}(A)$, is $\max\{\text{branch}_A(w) \mid w \in L(A)\}$ if this quantity is defined, and otherwise ∞ .

Hromkovic et al. [8] define three measures non-determinism for a finite automaton A : *advice*(A), *computations*(A), and *ambig*(A). These measures are defined as follows: *advice*(A) is the maximum number of non-deterministic choices during any computation of A , *computations*(A) is the maximum number of different computations of A on any word,³ and *ambig*(A) is the maximum number of different *accepting* computation of A on any word. For the formal definitions of these concepts, we refer to [8].

2.2 A Notion of Very Little Non-Determinism

Next we define the notion of a δ NFA. The intuition is that such an automaton should allow only a very small amount of non-determinism.

Definition 1. A δ NFA is an NFA A with the following properties

- A is unambiguous;
- $\text{branch}(A) \leq 2$; and
- there is at most one pair (q, a) such that $\text{degree}(q, a) = 2$.

For δ NFAs, we have that $\text{degree}(A) \leq 2$, $\text{advice}(A) \leq 1$, $\text{computations}(A) \leq 2$, and $\text{ambig}(A) = 1$. Notice that any one of $\text{degree}(A) = 1$, $\text{advice}(A) = 0$, or $\text{computations}(A) = 1$ implies that A is deterministic. Also, $\text{ambig}(A) = 1$ is the minimum value possible for any automaton that accepts at least one string.

2.3 The Minimization Problem

We define the minimization problem in two flavors. For two classes of finite automata \mathcal{N}_1 and \mathcal{N}_2 the $\mathcal{N}_1 \rightarrow \mathcal{N}_2$ *minimization problem* is the following problem. Given a finite automaton A in \mathcal{N}_1 and an integer k , does there exist a finite automaton B in \mathcal{N}_2 of size at most k such that $L(A) = L(B)$? For a class \mathcal{N} of finite automata, the *minimization problem for \mathcal{N}* is then simply the $\mathcal{N} \rightarrow \mathcal{N}$ minimization problem.

Suppose that the $\mathcal{N}_1 \rightarrow \mathcal{N}_2$ minimization problem is hard for a complexity class C , and let \mathcal{N}_3 be a class of automata such that $\mathcal{N}_1 \subseteq \mathcal{N}_3$. Then the $\mathcal{N}_3 \rightarrow \mathcal{N}_2$ minimization problem is also trivially hard for C . However, assuming that $\mathcal{N}_1 \rightarrow \mathcal{N}_2$ is hard for C and that $\mathcal{N}_2 \subseteq \mathcal{N}_3$, there is, as far as we know, no general argument that also makes the $\mathcal{N}_1 \rightarrow \mathcal{N}_3$ minimization problem hard for C , as finding a small \mathcal{N}_3 automaton might be easier than finding a small \mathcal{N}_2

³ Hromkovic et al. wrote $\text{leaf}(A)$ instead of $\text{computations}(A)$.

automaton in general.⁴ Therefore, we will prove directly that minimization is NP-hard for *all* classes of automata between δ NFAs and NFAs.

2.4 Are δ NFAs the Closest Possible to Determinism?

Before we give more intuition about this question, we first note that there are in fact *two* incomparable notions of determinism for finite automata: determinism and *reverse* determinism.⁵ Both classes can be efficiently minimized by the same algorithm, modulo a simple pre- and post-processing step for reverse deterministic automata. We view these two classes as the two possible “optima” in the spectrum of determinism, as they arise very naturally from the fact that one can either read strings from left to right or from right to left. From now on, we only consider the proximity of δ NFAs to (left-to-right deterministic) DFAs.

We believe that one can always think of more and more exotic notions of non-determinism that come closer and closer to DFAs.⁶ We provide an example here. Define the class \mathcal{C} to be the class of δ NFAs with the additional condition that, for each word w , there can be at most one *rejecting* computation of A on w . (Thus, for each w , there can be at most two runs — one accepting and one rejecting.) This notion of non-determinism lies strictly between DFAs and δ NFAs ($\text{DFA} \subseteq \mathcal{C} \subseteq \delta\text{NFA}$).

Consider the minimization problem for \mathcal{C} and let A be an arbitrary automaton in \mathcal{C} . We will argue that the minimal \mathcal{C} -automaton for $L(A)$ is a DFA. Suppose that A is not a DFA. Let q and a be the unique state and label so that $\text{degree}(q, a) = 2$. Let q_1 and q_2 be the two states such that $q \xrightarrow{a} q_1$ and $q \xrightarrow{a} q_2$ are in $\text{Rules}(A)$. Let w be an arbitrary string that leads A to state q . By definition of \mathcal{C} , A must accept every string of the form waw' , where w' is an arbitrary word in $\text{Alpha}(A)^*$. (If waw' would be rejected, then there would be two rejecting runs, one over q_1 and one over q_2 .) Therefore, we can make A strictly smaller by merging the two states q_1, q_2 to q_3 , removing all outgoing transitions from q_3 , making q_3 a final state, and adding loop transitions from q_3 to itself for each alphabet symbol. Moreover, by this operation, A becomes deterministic. Hence, every automaton A in \mathcal{C} that is not a DFA can be rewritten as a smaller DFA. This means that, in the class \mathcal{C} , the minimal automata are the DFAs. In particular, this also puts the minimization problem for \mathcal{C} into PTIME.

From the above it is clear that δ NFAs are certainly not the closest possible to determinism that one can get. Rather, it is the closest class to DFAs we were able to find that takes advantage of the succinctness of nondeterminism in a nontrivial way.

Our NP-hardness result for the minimization of δ NFAs therefore puts the tractability frontier precisely between δ NFAs and the above mentioned class \mathcal{C} ; two classes that are extremely close to one another.

⁴ This is also why, e.g., Malcher explicitly proves NP-hardness for minimizing various classes of automata that are included in one another (Lemmas 3 and 11 in [12]).

⁵ The latter is the class for which the inverted transitions are deterministic.

⁶ One could, for instance, take the class of DFAs and add a single NFA.

3 Minimizing Non-Deterministic Automata is Hard

The main result of this section is the following

Theorem 2. *Let \mathcal{N} be a class of finite automata. If $\delta\text{NFA} \subseteq \mathcal{N}$ then $\text{DFA} \rightarrow \mathcal{N}$ minimization is NP-hard.*

Corollary 3. *For each class \mathcal{N} of finite automata such that $\delta\text{NFA} \subseteq \mathcal{N}$, the minimization problem for \mathcal{N} is NP-hard.*

We start by formally defining the decision problems that are of interest to us, and then sketch an intuitive overview of our proof. Given an undirected graph $G = (V, E)$ such that V is its set of vertices and $E \subseteq V \times V$ is its set of edges, we say that a set of vertices $VC \subseteq V$ is a *vertex cover* of G if, for every edge $(v_1, v_2) \in E$, VC contains v_1 , v_2 , or both.

If B and C are finite collections of finite sets, we say that B is a *set basis* for C if, for each $c \in C$, there is a subcollection B_c of B whose union is c . We say that B is a *normal set basis* for C if, for each $c \in C$, there is a *pairwise disjoint* subcollection B_c of B whose union is c . We say that B is a *separable normal set basis* for C if B is a normal set basis for C and B can be written as a disjoint union $B_1 \uplus B_2$ such that, for each $c \in C$, the subcollection B_c of B contains at most one element from B_1 and at most one from B_2 .

The following decision problems are considered in this paper. *Vertex Cover* asks, given a pair (G, k) where G is a graph and k is an integer, whether there exists a vertex cover of G of size at most k . *Set Basis*, *Normal Set Basis*, and *Separable Normal Set Basis* ask, given a pair (C, s) where C is a finite collection of finite sets and s is an integer, whether there exists a set basis, resp., normal set basis, resp., separable normal set basis for C containing at most s sets.

The proof of Theorem 2 proceeds in several steps. First, we provide a slightly modified version of a known reduction from Vertex Cover to Normal Set Basis (Lemma 4 in [11]), showing that the latter problem is NP-hard. Second, we proceed to show that the set \mathbf{I} of instances of Normal Set Basis obtained through this reduction has a number of interesting properties (Lemma 5). In particular, we show that if such an instance has a set basis of a certain size s , then it also has a *normal set basis* of size s . Third, we show that the Normal Set Basis problem, for instances in \mathbf{I} reduces to minimization for δNFAs (Lemma 6).

The statement of Theorem 2 says that given a DFA, finding the minimal equivalent automaton in class \mathcal{N} is NP-hard, for any class of finite automata that contains the δNFAs . As argued in Section 2.3, using a DFA instead of a δNFA as input of the problem strengthens the statement. Also, showing that $\text{DFA} \rightarrow \delta\text{NFA}$ is NP-hard doesn't immediately imply that $\text{DFA} \rightarrow \mathcal{N}$ is hard for every \mathcal{N} that contains all δNFAs . To show that this is actually the case, we prove that for the languages obtained in our reduction, the minimal NFAs are precisely one state smaller than the minimal δNFAs (Lemma 6). For these languages, the minimization problem for δNFAs and for NFAs is essentially the same problem.

We revisit a slightly modified reduction which is due to Jiang and Ravikumar [11], as our further results rely on a construction in their proof.

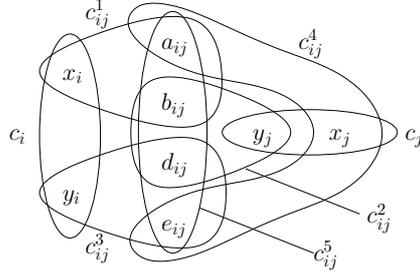


Fig. 1. The constructed sets $c_i, c_j, c_{ij}^1, \dots, c_{ij}^5$ in the proof of Lemma 4.

Lemma 4 (Jiang and Ravikumar [11]). *Normal Set Basis is NP-complete.*

Proof (Sketch). Obviously, Normal Set Basis is in NP. Indeed, given an input (C, s) for Normal Set Basis, the NP algorithm simply guesses a collection B containing at most s sets, guesses the subcollections B_c for each $c \in C$, and verifies whether the sets B_c satisfy the necessary conditions.

We give a reduction from Vertex Cover to Normal Set Basis but omit the correctness proof. Given an input (G, k) of Vertex Cover, where $G = (V, E)$ is a graph and k is an integer, we construct in LOGSPACE an input (C, s) of Normal Set Basis, where C is a finite collection of finite sets and s is an integer. In particular, (C, s) is constructed such that G has a vertex cover of size at most k if and only if C has a normal set basis containing at most s sets.

For a technical reason which will become clear in later proofs, we assume without loss of generality that $k < |E| - 3$. Notice that, under this restriction, Vertex Cover is still NP-complete under LOGSPACE reductions. Indeed, if $k \geq |E| - 3$, Vertex Cover can be solved in LOGSPACE by testing all possibilities of the at most 3 vertices which are not in the vertex cover, and verifying that there does not exist an edge between 2 of these 3 vertices.

Formally, let $V = \{v_1, \dots, v_n\}$. For each $i = 1, \dots, n$, define c_i to be the set $\{x_i, y_i\}$ which intuitively corresponds to the node v_i . Let (v_i, v_j) be in E with $i < j$. To each such edge we associate five sets as follows:

$$\begin{aligned} c_{ij}^1 &:= \{x_i, a_{ij}, b_{ij}\}, & c_{ij}^4 &:= \{x_j, a_{ij}, e_{ij}\}, \text{ and} \\ c_{ij}^2 &:= \{y_j, b_{ij}, d_{ij}\}, & c_{ij}^5 &:= \{a_{ij}, b_{ij}, d_{ij}, e_{ij}\}. \\ c_{ij}^3 &:= \{y_i, d_{ij}, e_{ij}\}, \end{aligned}$$

Figure 1 contains a graphical representation of the constructed sets $c_i, c_j, c_{ij}^1, \dots, c_{ij}^5$ for some $(v_i, v_j) \in E$. Then, define

$$C := \{c_i \mid 1 \leq i \leq n\} \cup \{c_{ij}^t \mid (v_i, v_j) \in E, i < j, \text{ and } 1 \leq t \leq 5\}$$

and $s := n + 4|E| + k$. Notice that the collection C contains $n + 5|E|$ sets. Obviously, C and s can be constructed from G and k in LOGSPACE.

We show that the reduction is also correct, that is, that G has a vertex cover of size at most k if and only if C has a (separable) normal set basis containing at most s sets.

(\Rightarrow): Let G have a vertex cover VC of size k . We need to show that C has a normal set basis B containing at most $s = n + 4|E| + k$ sets.

To this end, we define a collection B of sets as follows. For every $v_i \in V$,

- if $v_i \in VC$, we include both $\{x_i\}$ and $\{y_i\}$ in B ;
- otherwise, we include $c_i = \{x_i, y_i\}$ in B .

The number of sets included in B so far is $2k + (n - k) = k + n$. Let $e = (v_i, v_j)$ (where $i < j$) be an arbitrary edge in G . Since VC is a vertex cover, either v_i or v_j (or both) is in VC . When v_i is in VC , we additionally include the sets

$$\begin{aligned} r_{ij}^1 &:= \{a_{ij}, b_{ij}\}, & r_{ij}^2 &:= \{d_{ij}, e_{ij}\}, \\ r_{ij}^3 &:= \{y_j, b_{ij}, d_{ij}\}, & \text{and } r_{ij}^4 &:= \{x_j, a_{ij}, e_{ij}\} \end{aligned}$$

in B . When v_i is not in VC , we additionally include the sets

$$\begin{aligned} r_{ij}^5 &:= \{a_{ij}, e_{ij}\}, & r_{ij}^6 &:= \{b_{ij}, d_{ij}\}, \\ r_{ij}^7 &:= \{x_i, a_{ij}, b_{ij}\}, & \text{and } r_{ij}^8 &:= \{y_i, d_{ij}, e_{ij}\} \end{aligned}$$

in B . This completes the definition of B . Notice that, when $v_i \in VC$, c_{ij}^1 , c_{ij}^3 , and c_{ij}^5 can be expressed as a disjoint union of members of B as

$$c_{ij}^1 = \{x_i\} \uplus r_{ij}^1, \quad c_{ij}^3 = \{y_i\} \uplus r_{ij}^2, \quad c_{ij}^5 = r_{ij}^1 \uplus r_{ij}^2$$

and that $c_{ij}^2 = r_{ij}^3$ and $c_{ij}^4 = r_{ij}^4$ are members of B . Analogously, when $v_i \notin VC$, c_{ij}^2 , c_{ij}^4 , and c_{ij}^5 can be expressed as a disjoint union of members of B as

$$c_{ij}^2 = \{y_j\} \uplus r_{ij}^6, \quad c_{ij}^4 = \{x_j\} \uplus r_{ij}^5, \quad c_{ij}^5 = r_{ij}^5 \uplus r_{ij}^6$$

and $c_{ij}^1 = r_{ij}^7$ and $c_{ij}^3 = r_{ij}^8$ are members of B . Since the total number of sets included in B for each edge is four, B contains $(k + n) + 4|E| = s$ sets. From the above argument it is also obvious that B is a normal set basis for C .

Notice that B is in fact a separable normal set basis for C . Indeed, we can partition B into the sets

$$\begin{aligned} B_1 = & \{ \{x_i\}, \{x_j, y_j\} \mid v_i \in VC, v_j \notin VC \} \\ & \cup \{ r_{ij}^2, r_{ij}^3 \mid (v_i, v_j) \in E, i < j, v_i \in VC \} \\ & \cup \{ r_{ij}^6, r_{ij}^7 \mid (v_i, v_j) \in E, i < j, v_i \notin VC \} \end{aligned}$$

and

$$\begin{aligned} B_2 = & \{ \{y_i\} \mid v_i \in VC \} \\ & \cup \{ r_{ij}^1, r_{ij}^4 \mid (v_i, v_j) \in E, i < j, v_i \in VC \} \\ & \cup \{ r_{ij}^5, r_{ij}^8 \mid (v_i, v_j) \in E, i < j, v_i \notin VC \}, \end{aligned}$$

which satisfy the necessary condition.

(\Leftarrow): Suppose that C has a normal set basis B containing at most $s = n + 4|E| + k$ sets. We can assume without loss of generality that no proper subcollection of B is a normal set basis. We show that G has a vertex cover VC of size at most k . Define $VC = \{v_i \mid \text{both } \{x_i\} \text{ and } \{y_i\} \text{ are in } B\}$. Let k' be the number of elements in VC . The number of sets in B consisting of only x_i and/or y_i is at least $n + k'$. This can be seen from the fact that B must have the subset c_i for all i such that $v_i \notin VC$. Thus, there are $n - k'$ such sets in addition to $2k'$ singleton sets corresponding to i 's such that $v_i \in VC$. Let $E' \subseteq E$ be the set of edges covered by VC , that is, $E' = \{(v_i, v_j) \mid v_i \text{ or } v_j \text{ is in } VC\}$. The following observation can easily be shown (by checking all possibilities):

Observation: For any $e \in E'$ at least four sets of B (excluding sets $c_i, c_j, \{x_i\}, \{y_i\}, \{x_j\}$, or $\{x_j\}$) are necessary to be a normal set basis for the five sets $c_{ij}^t, t = 1, \dots, 5$. Further, at least five sets (excluding sets $c_i, c_j, \{x_i\}, \{y_i\}, \{x_j\}$, or $\{x_j\}$) are required to be a normal set basis for them if $e \notin E'$. Notice that, for $e \notin E'$, $\{x_i\}$ and $\{y_i\}$, or $\{x_j\}$ and $\{y_j\}$ are never both in B , by definition of E' .

Now the total number of sets needed to cover C is at least $n + k' + 4|E'| + 5(|E| - |E'|)$, which we know is at most $s = n + 4|E| + k$. Hence, we obtain that $n + k' + 5|E| - |E'| \leq n + 4|E| + k$, which implies that $k' + |E| - |E'| \leq k$. We conclude the proof by showing that there is a vertex cover VC' of size $|E| - |E'| + k'$. Add one of the end vertices of each edge $e \in E - E'$ to VC . This vertex cover is of size $|E| - |E'| + k' \leq k$. \square

The next lemma now follows from the proof of Lemma 4. It defines a set of inputs \mathbf{I} for which Normal Set Basis is NP-complete and further shows that for any $(C, s) \in \mathbf{I}$, the collection C has a set basis of size s if and only if C also has a separable normal set basis of size s . Of course, the latter property does not hold for the set of all possible inputs for the normal set basis problem.

Lemma 5. *There exists a set of inputs \mathbf{I} for Normal Set Basis, such that*

- (1) *Normal Set Basis is NP-complete for inputs in \mathbf{I} ;*
- (2) *for each (C, s) in \mathbf{I} , C contains every set at most once and $s < |C| - 3$;*
- (3) *for each $(C, s) \in \mathbf{I}$, the following are equivalent:*
 - (a) *C has a set basis containing s sets.*
 - (b) *C has a separable normal set basis containing s sets.*
- (4) *for each (C, s) in \mathbf{I} , each solution B for (C, s) writes at least two sets of C as a union of at least two sets in B .*

Proof. The set \mathbf{I} is obtained by applying the reduction in Lemma 4 to inputs (G, k) of Vertex Cover such that $k \leq |E| - 3$. This immediately shows (1) and (2). We continue by proving the other cases.

(3) The direction from (b) to (a) is trivial. For the other direction the full proof of Lemma 4 actually shows that if G has a vertex cover of size k , then C has a *separable* normal set basis containing s sets. Conversely, if C has a *normal* set basis containing at most s sets, then G has a vertex cover of size k . This implies that C also has a separable normal set basis containing s sets.

Hence, we still need to prove that, if C has a set basis of at most s sets, then C also has a normal set basis containing at most s sets. Let (C, s) be an instance in \mathbf{I} , i.e., there is an $n \in \mathbb{N}$ and $E \subseteq \{(i, j) \mid 1 \leq i < j \leq n\}$ such that $C = \{c_i \mid 1 \leq i \leq n\} \cup \{c_{ij}^r \mid (i, j) \in E \wedge 1 \leq r \leq 5\}$, and suppose C has a set basis $B = \{b_1, \dots, b_s\}$ of size s . We construct a *normal* set basis for C of size s .

To this end, we will show a sequence of assumptions that we can make about B without loss of generality. These assumptions, put together, will imply that B is a *normal* set basis for C , and thus show that if there is a set basis of size s , then there is also a normal set basis of size s .

Suppose that there is an i such that B contains both $\{x_i\}$ and $\{x_i, y_i\}$. Then we can replace $\{x_i, y_i\}$ with $\{y_i\}$ and still have a set basis for C , since c_i is the only set in C of which $\{x_i, y_i\}$ is a subset. This gives us our first assumption.

Assumption 1. For every $i \in \{1, \dots, n\}$, set basis B either contains $\{x_i, y_i\}$ or $\{x_i\}$ and $\{y_i\}$, but never both $\{x_i, y_i\}$ and $\{x_i\}$ (or, symmetrically, $\{x_i, y_i\}$ and $\{y_i\}$).

Suppose there are $1 \leq i < j \leq n$ and $1 \leq r \leq 4$ such that c_{ij}^r cannot be formed as a disjoint union of sets from B . Without loss of generality, we may assume that $r = 1$, i.e., $c_{ij}^1 = \{x_i, a_{ij}, b_{ij}\}$, since all other cases follow by symmetry. Since there are no disjoint sets from B whose union is c_{ij}^1 , there must be two different sets b^1 and b^2 in B that are subsets of c_{ij}^1 and contain precisely two elements each. At least one of these subsets must contain x_i . Assume w.l.o.g. that this set is b^1 . No subset of size two of c_{ij}^1 that contains x_i is a subset of any set of C other than c_{ij}^1 . This means that we can replace b^1 with $b^1 - b^2$ in B and still have a set basis of size at most s , and gives us our second assumption.

Assumption 2. For any i, j and any $r \in \{1, \dots, 4\}$, the set c_{ij}^r can be formed as a union of disjoint sets from B .

If B satisfies Assumptions 1 and 2, but is not normal, there are $1 \leq i < j \leq n$ such that c_{ij}^5 cannot be formed as a disjoint union of sets from B . Let B_{ij} be a subset of B such that the union of the sets in B_{ij} is c_{ij}^5 . We can assume that B_{ij} is inclusion free, i.e., there are no two sets in B_{ij} such that one is a subset of the other, since if there are $b^1, b^2 \in B_{ij}$ with $b^1 \subset b^2$, we can replace b^2 with $b^2 - b^1$.

Assumption 3. The collection B_{ij} is inclusion free.

If B_{ij} has four members, then we can replace B_{ij} with the four singletons and still have a set basis of size s .

Assumption 4. The collection B_{ij} has at most three members.

Suppose there is a set b in B_{ij} such that b is not a subset of any set in C other than c_{ij}^5 . Then we can replace b with $b - (\bigcup_{b' \in B_{ij} - \{b\}} b')$ in B and still have a set basis of size at most s .

Assumption 5. Each member of B_{ij} is a subset of some set from C other than c_{ij}^5 . In particular, this means that each member of B_{ij} has at most two elements.

If we take three different subsets of c_{ij}^5 , with at most two elements, that are also subsets of other sets from C than c_{ij}^5 , then at least two of them are disjoint; see Figure 1. Let these two disjoint sets be b^1 and b^2 . If b^1 and b^2 both contain two elements, we can replace B_{ij} with $\{b^1, b^2\}$.

Assumption 6. There are at most two sets in B_{ij} with two elements. Furthermore, these two sets must overlap. This means that B_{ij} must have exactly three members, two with two elements and one singleton. Without loss of generality, we may assume that $B_{ij} = \{\{a_{ij}, b_{ij}\}, \{b_{ij}, d_{ij}\}, \{e_{ij}\}\}$. (All other cases are symmetrical.) We may also assume that neither $\{a_{ij}\}$ nor $\{b_{ij}\}$ belong to B . (If $\{a_{ij}\}$ belongs to B then we can replace $\{a_{ij}, b_{ij}\}$ by $\{b_{ij}\}$ in B_{ij} , and if $\{b_{ij}\}$ belongs to B we can replace $\{b_{ij}, d_{ij}\}$ by $\{d_{ij}\}$ in B_{ij} .)

If B_{ij} satisfies Assumption 6, in order to form c_{ij}^1 either $\{x_i\}$, $\{x_i, a_{ij}\}$, $\{x_i, b_{ij}\}$, or $\{x_i, a_{ij}, b_{ij}\}$ must be a member of B . If it is not $\{x_i\}$, we can replace it with $\{x_i\}$, since none of the other sets is a subset of any other set in C than c_{ij}^1 . But if $\{x_i\} \in B$ we can also assume that $\{y_i\} \in B$. To form c_{ij}^3 , B must, apart from $\{y_i\}$ and $\{e_{ij}\}$, contain some subset of c_{ij}^3 that contains d_{ij} . Since we have both $\{y_i\}$ and $\{e_{ij}\}$ in B , we may replace this subset with $\{d_{ij}\}$. Once we have $\{d_{ij}\}$ in B , we can replace $\{b_{ij}, d_{ij}\}$ by $\{b_{ij}\}$ in B .

Assumption 7. There are disjoint members of B whose union is c_{ij}^5 .

Together, Assumptions 1,2, and 7 imply that B is in fact a normal set basis for C . Since each assumption was made without loss of generality, we have shown that from any set basis for C of size s we can form a *normal* set basis for C of size at most s .

(4) We simply observe that a normal set basis writing at most one set of C as a union of at least two sets must contain at least $|C|$ sets, and hence cannot be a solution for (C, s) , since $s < |C|$. \square

The proof of the following lemma is partly inspired by the proof of Theorem 3.1 of [11], but we significantly strengthen it for our purposes.

Lemma 6. *There exists a set of regular languages \mathcal{L} such that*

- (1) *DFA \rightarrow δ NFA minimization is NP-complete for DFAs accepting \mathcal{L} and*
- (2) *for each $L \in \mathcal{L}$, the size of the minimal NFA for L is equal to the size of the minimal δ NFA for L , minus 1.*

Proof. The NP upper bound is immediate, as equivalence testing for unambiguous finite automata is in PTIME [17]. We can guess a δ NFA of sufficiently small size and test in PTIME whether it is equivalent to the given DFA.

For the lower bound, we reduce from Separable Normal Set Basis. To this end, let (C, s) be an input of Separable Normal Set Basis. Hence, C is a collection of n sets and s is an integer. According to Lemma 5, we can assume without loss of generality that $(C, s) \in \mathbf{I}$, that is, C has a separable normal set basis containing s sets if and only if C has a normal set basis of size s . Moreover, we can assume that $s < n - 3$.

We construct in LOGSPACE a DFA A and an integer ℓ such that the following are equivalent:

- C has a separable normal set basis of size at most s .
- There exists a δ NFA N_δ for $L(A)$ of size at most ℓ .
- There exists an NFA N for $L(A)$ of size at most $\ell - 1$.

The DFA A accepts the language $\{acb \mid c \in C \text{ and } b \in c\}$, which is a finite language of strings of length three.

Formally, let $C = \{c_1, \dots, c_n\}$ and $c_i = \{b_{i,1}, \dots, b_{i,n_i}\}$. Then, A is defined over $\text{Alpha}(A) = \{a\} \cup \bigcup_{1 \leq i \leq n} \{c_i, b_{i,1}, \dots, b_{i,n_i}\}$. The state set of A is $\text{States}(A) = \{q_0, q'_0, q_1, \dots, q_n, q_f\}$, and the initial and final state sets of A are q_0 and q_f , respectively. The transitions $\text{Rules}(A)$ are formally defined as follows:

- $q_0 \xrightarrow{a} q'_0$;
- for every $i = 1, \dots, n$, $q'_0 \xrightarrow{c_i} q_i$; and
- for every $i = 1, \dots, n$ and $j = 1, \dots, n_i$, $q_i \xrightarrow{b_{i,j}} q_f$.

Finally, define $\ell := s + 4$.

Obviously, A and ℓ can be constructed from C and s using logarithmic space. Observe that due to Lemma 5, C contains every set at most once, and hence does not contain $c_i = c_j$ with $i \neq j$. Hence, A is a minimal DFA for $L(A)$.

We now show that,

- (a) if C has a separable normal set basis containing at most s sets, then there exists a δ NFA N_δ for $L(A)$ of size at most ℓ and an NFA N for $L(A)$ of size at most $\ell - 1$;
- (b) if there exists a δ NFA N_δ for $L(A)$ of size at most ℓ then C has a separable normal set basis containing at most s sets; and
- (c) if there exists an NFA N for $L(A)$ of size at most $\ell - 1$ then C has a separable normal set basis containing at most s sets.

(a) Assume that C has a separable normal set basis containing s sets. We construct a δ NFA N_δ for $L(A)$ of size $\ell = s + 4$.

Let $B = \{r_1, \dots, r_s\}$ be the separable normal set basis for C containing s sets. Also, let B_1 and B_2 be disjoint subcollections of B such that each element of C is either an element of B_1 , an element of B_2 , or a disjoint union of an element of B_1 and an element of B_2 .

To describe N_δ , we first fix the representation of each set c in C as a disjoint union of at most one set in B_1 and at most one set in B_2 . Say that each basic member of B in this representation *belongs to* c .

We define the state set of N_δ as $\text{States}(N_\delta) = \{q_0, q_1, q_2, q_f\} \cup \{r_i \in B_1\} \cup \{r_i \in B_2\}$. The transition rules of N_δ are defined as follows. First, $\text{Rules}(N_\delta)$ contains the non-deterministic transitions $q_0 \xrightarrow{a} q_1$ and $q_0 \xrightarrow{a} q_2$. Furthermore, for every $i = 1, \dots, n$, $j = 1, \dots, s$, and $m = 1, 2$, $\text{Rules}(N_\delta)$ contains the rule

- $q_m \xrightarrow{c_i} r_j$, if $r_j \in B_m$ and r_j belongs to c_i ; and
- $r_j \xrightarrow{b} q_f$, if $b \in r_j$.

Notice that the size of N_δ is $|B| + 4 = s + 4 = \ell$. By construction, we have that $L(N_\delta) = L(A)$.

We argue that N_δ is a δ NFA. First, we argue that the transitions $q_0 \xrightarrow{a} q_1$ and $q_0 \xrightarrow{a} q_2$ are the only non-deterministic transitions. By construction, all transitions going to q_f are deterministic. It is also easy to see that all outgoing transitions from q_1 are deterministic because, if we assume that N_δ contains transitions of the form $q_1 \xrightarrow{c_i} r_j$ and $q_1 \xrightarrow{c_i} r_{j'}$ with $j \neq j'$, this would mean that both r_j and $r_{j'}$ belong to c_i , which contradicts the definition of B_1 . The argument for q_2 is analogous.

Next, we show that N_δ is unambiguous. Towards a contradiction, assume that the string $ac_i b$ has two accepting runs. The only way in which this can happen is that one run visits state q_1 and the other run visits state q_2 . Let r_{j_1} (respectively, r_{j_2}) be the state such that $q_1 \xrightarrow{c_i} r_{j_1}$ (respectively, $q_2 \xrightarrow{c_i} r_{j_2}$) are transitions in N_δ . By construction of N_δ , we have that $j_1 \neq j_2$. But this means that both r_{j_1} and r_{j_2} belong to c_i , and their intersection contains the element b , which contradicts the disjointness condition of the normal set basis B .

Finally, the NFA N for $L(A)$ is obtained by merging the two states q_1 and q_2 from N_δ . Obviously, the size of N is $\ell - 1$.

(b) Assume that $L(A)$ can be accepted by a δ NFA N_δ of size at most ℓ . We can assume that N_δ is minimal. We need to show that there exists a separable normal set basis for C containing at most $s = \ell - 4$ sets.

Recall that we assumed that $s < n - 3$. Hence, we have that $\ell = s + 4 < n + 1 = |A| - 1$. As we observed that A is a minimal DFA for $L(A)$, it must be the case that N_δ has at least one non-deterministic transition.

Notice that, as N_δ is minimal and accepts only strings of length three, we can assume that N_δ has a unique final state. Let q_0 and q_f be the initial and final state of N_δ , respectively. Furthermore, we can partition N_δ 's states into four sets \mathcal{Q}_0 , \mathcal{Q}_1 , \mathcal{Q}_2 , and \mathcal{Q}_3 such that, for each $0 \leq i \leq 3$, \mathcal{Q}_i is precisely the set of states that N_δ can be in after having read a string of length i . For each state $q \in \mathcal{Q}_2$, define a set $B_q = \{b \mid q \xrightarrow{b} q_f \in \text{Rules}(N_\delta)\}$.

Next, we show that the collection $B = \{B_q \mid q \in \mathcal{Q}_2\}$ is a normal set basis for C . By definition of $L(A)$, we have that every $c \in C$ is the union of $B_c := \{B_q \mid \exists p \in \mathcal{Q}_1 : p \xrightarrow{c} q \in \text{Rules}(M_i)\}$. It remains to show that B_c is also a disjoint subcollection of B . When B_c contains only one set, there is nothing to prove. Towards a contradiction, assume that B_c contains two different sets B_{q_1} and B_{q_2} such that $a \in B_{q_1} \cap B_{q_2}$. By definition of B , this would mean that the string acb has two accepting runs: qq_1q_f and qq_2q_f with $q_1 \neq q_2$. But as N_δ is unambiguous, this is impossible. Hence, B_c is a disjoint subcollection of B .

Finally, we want to prove that B contains at most $\ell - 4$ sets. Notice that B contains at most $\ell - 4$ sets if and only if \mathcal{Q}_1 contains at least two states. Towards a contradiction, assume that \mathcal{Q}_1 is a singleton q . By Lemma 5, we know that there are at least two sets c_1, c_2 such that B_{c_1} and B_{c_2} contain two sets. In other words, there are at least two alphabet symbols c_1 and c_2 that have non-deterministic outgoing transitions from q , which contradicts that N_δ is a δ NFA. Hence, B is a normal set basis for C of size at most $\ell - 4$. As $(C, s) \in \mathbf{I}$, we have that C has a separable normal set basis of size $s = \ell - 4$ by Lemma 5.

(c) Let N be an NFA for $L(A)$ of size at most $\ell - 1$. We can assume that N is minimal. We will show that there exists a set basis for C containing at most $s = \ell - 4$ sets.

Recall that we assumed that $s < n - 3$. Hence, we have that $\ell = s + 4 < n + 1 = |A| - 1$. As we observed that A is a minimal DFA for $L(A)$, it must be the case that N has at least one non-deterministic transition.

Notice that, as N is minimal and accepts only strings of length three, we can assume that N has a unique final state. Let q_0 and q_f be the initial and final state of N , respectively. Furthermore, we can partition N 's states into four sets $\mathcal{Q}_0, \mathcal{Q}_1, \mathcal{Q}_2$, and \mathcal{Q}_3 such that, for each $0 \leq i \leq 3$, \mathcal{Q}_i is precisely the set of states that N can be in after having read a string of length i . For each state $q \in \mathcal{Q}_2$, define a set $B_q = \{b \mid q \xrightarrow{b} q_f \in \text{Rules}(N_\delta)\}$. Next, we show that the collection $B = \{B_q \mid q \in \mathcal{Q}_2\}$ is a set basis for C . By definition of $L(A)$, we have that every $c \in C$ is the union of $B_c := \{B_q \mid \exists p \in \mathcal{Q}_1 : p \xrightarrow{c} q \in \text{Rules}(M_i)\}$. As $\mathcal{Q}_0, \mathcal{Q}_1$, and \mathcal{Q}_3 contain at least one state, we also have that B contains at most $\ell - 1 - 3 = s$ sets.

From Lemma 5, it now follows that C also has a separable normal set basis containing at most s sets. \square

We are now ready to finish the proof of Theorem 2.

Proof of Theorem 2. Let \mathcal{N} be a class of finite automata. If $\delta\text{NFA} \subseteq \mathcal{N}$ then $\text{DFA} \rightarrow \mathcal{N}$ minimization is NP-hard.

Proof. In this Section we provided a reduction from Vertex Cover to $\text{DFA} \rightarrow \delta\text{NFA}$ minimization, and showed that, for the regular languages we consider, the minimal NFA is 1 state smaller than the minimal δNFA .

Let \mathcal{N} be a class of finite automata such that $\delta\text{NFA} \subseteq \mathcal{N} \subseteq \text{NFA}$. As was shown in the proof of Lemma 6, any decision algorithm for $\text{DFA} \rightarrow \mathcal{N}$ minimization can approximate the $\text{DFA} \rightarrow \delta\text{NFA}$ minimization problem within a term 1 (as the minimal NFA is only one state smaller than the minimal δNFA).

As can be seen from the other proofs in Section 3, this approximation algorithm for $\text{DFA} \rightarrow \delta\text{NFA}$ minimization can easily be adapted to an approximation algorithm for Vertex Cover within a term 1. As we know that it is NP-hard to approximate Vertex Cover within a constant term, we can conclude that $\text{DFA} \rightarrow \mathcal{N}$ minimization is also NP-hard. \square

Until now, our results focused on classes of finite automata that can accept all regular languages. Our proof shows that this is not even necessary, as the

NP-hard instances we construct only accept strings of length tree. Therefore, we also have the following Corollary.

Corollary 7. *Let δNFA_{finite} be the class of δNFA s that accept only finite languages. Let \mathcal{N} be class of finite automata. If $\delta NFA_{finite} \subseteq \mathcal{N}$ then the $DFA \rightarrow \mathcal{N}$ minimization problem is NP-hard.*

4 Succinctness and Uniqueness

As mentioned in the introduction, when a developer selects a description mechanism for regular languages, she faces a trade-off between succinctness and complexity of minimization. The following proposition shows that in the case of δNFA s, the succinctness bought at the price of NP-completeness is limited.

Proposition 8. *For every δNFA of size n , there is an equivalent DFA of size $O(n^2)$.*

Proof. For every finite automaton that can choose non-deterministically between two different start states but is otherwise deterministic (a 2-MDFA), there is an equivalent DFA with at most quadratically many states [7].

Let A be a δNFA , and let (q, a) be the only pair in $States(A) \times Alpha(A)$ such that $degree(q, a) = 2$. Let q_1 and q_2 be the two states reachable from q when reading an a . If we remove all states from A that are reachable neither from q_1 nor from q_2 , and make q_1 and q_2 initial states, we obtain a 2-MDFA A' . We can compute the smallest DFA A'' equivalent to A' in quadratic time. Now, all we need to do is to add A'' to A and replace the two rules going from q to q_1 and q_2 , respectively, by a single rule that reads an a and goes to the initial state of A'' . The size of the new, deterministic, automaton is $|States(A)| + |States(A'')| = O(|States(A)|^2)$. \square

On the other hand, if we were to remove the $branch(A) \leq 2$ condition in the definition of δNFA s, then there would be an exponential gain in succinctness. This is witnessed by the standard family of languages $(a+b)^*a(a+b)^n$ for $n \geq 0$ that shows that NFA s are exponentially more succinct than DFA s in general. The canonical NFA for this language is unambiguous and has only one pair (q, a) for which $degree(q, a) = 2$.

Proposition 9. *The minimal δNFA for a regular language is not unique.*

Proof. Consider the language L defined by the regular expression

$$r = (a+b)aaa + b(a+b)(a+b)b.$$

Figure 2 depicts two δNFA s, A and B , that both accept L . We argue that eight is the minimal number of states for any δNFA that accepts L . First, it is clear that any such automaton has to remember how many letters it has read so far. Second, the automaton has to have at least two states that can be reached after

reading one letter, one that accepts the string bbb and one that does not. Third, there must also be at least two states reachable after reading two letters, one that accepts the string bb and one that doesn't. Fourth, there must be at least two states that can be reached after reading three letters, one that accepts the string b and one that does not. Together with the fact that there has to be one initial state and at least one final state, this shows that any δ NFA for L needs at least eight states. \square

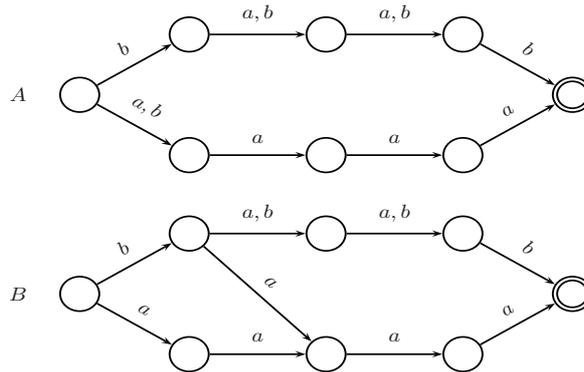


Fig. 2. Two minimal δ NFAs, that both accept the language L from the proof of Proposition 9.

5 Automata with Multiple Initial States

Throughout the paper, to simplify definitions, we have assumed that finite automata have a unique start state. As we mentioned in the Introduction, the minimization problem for finite automata that can non-deterministically choose between multiple initial states, but are otherwise deterministic, has also been studied [7, 12].

Proposition 10. *Minimization is NP-hard for unambiguous finite automata that have at most two initial states but are otherwise deterministic.*

Proof. We simply have to reconsider the proof of Lemma 6, change $L(A)$ so that it accepts the cb suffixes of its current language, consider the δ NFA we construct without its start state, and make the two successors of the start state initial states (and analogously for the NFA considered in the proof). The rest of the proof carries through virtually unchanged. \square

Together with Theorem 2 this answers all the open questions mentioned by Malcher [12].

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