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On the Complexity of Languages Definable by Hereditary Elementary Formal Systems

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Abstract

An elementary formal system (EFS) is a logical system that generates a language. In this paper, we consider a subclass of EFS, called hereditary EFS (HEFS). First, we compare HEFS with the following extensions of pattern languages: multi-pattern languages, languages defined by pattern grammars, and non-synchronized pattern languages. Particularly, we show that a subclass of HEFS called simple EFS (SEFS) precisely generates non-synchronized pattern languages, and the subclass of SEFS with only one predicate symbol precisely generates the languages defined by pattern grammars. Next, we analyze the complexity of the languages definable by HEFS. We show that HEFS exactly defines the complexity class P, the class of languages accepted by deterministic Turing machines in polynomial time. This seems to be the first result to characterize the class P by grammars, while various characterization results by automata, logic, recursive functions, algebraic systems, lambda calculus are shown in literatures. We also show that a subclass of HEFS, called linear HEFS, exactly defines NSPACE($\log n$). Finally, we consider the membership problem for HEFS, that is, the problem of, given a string w and an HEFS Γ, determining whether Γ generates w. We prove that the membership problem is EXPTIME-complete for HEFS, and NP-complete for SEFS.

1 Introduction

In this paper, we consider elementary formal systems (EFS's) which are logical systems introduced by Smullyan for the development of recursive function theory over strings [12]. An EFS is a set of definite clauses, called axioms, whose arguments have patterns. EFS's are extensively studied in formal language theory [2], logic programming [14], and computational learning theory [3, 10, 11]. A hereditary EFS (HEFS) is originally introduced by Miyano *et al.* to investigate the polynomial-time learnability of formal languages from examples [9]. An EFS is hereditary if, for each *axiom*

$A \leftarrow B_1, \ldots, B_m$, every pattern in $B_1, \ldots, B_m$ must appear as a subword of some pattern in A.

First, we study the simulation capability of a subclass of HEFS, called simple EFS (SEFS) defined by Arikawa [2]. The only relation among SEFS and other language classes known so far is the inclusion CFL $\subseteq$ SEFS $\subseteq$ CSL [3]. On the other hand, many extensions of CFL are proposed in literatures [8, 13]. These grammars generalize CFL, for example, by iterative substitutions for variables, or by parallel rewriting with productions. Non-synchronized pattern languages (NSPL) [8] are an example of the former type and extended OL-systems (EOL) [13] are an example of the latter type. Then, we show that the class NSPL is exactly the class of languages definable by SEFS. An interesting consequence of this result is that CFL, EOL, and NSPL have almost same space complexity modulo logspace reduction.

Next, we consider the computational complexity of the HEFS languages. Miyano *et al.* [9] showed that any HEFS language is decidable in polynomial time. Thus, the class HEFS is included in the complexity class P. We show that the converse is also true. Thus, P is the class of languages definable by HEFS. To prove this theorem, we simulate a two-way multihead alternating finite automaton [7] with an HEFS. In a sense, this result gives a framework for describing formal languages for which efficient parsers exist. As a consequence of this result, we have another subclass of HEFS, called linear HEFS, that precisely defines the class NSPACE($\log n$).

Finally, we investigate the computational complexity of the membership problem, which is the problem to, given a grammar G and a string w, decide whether G generates w. We show that the membership problem is EXPTIME-complete for HEFS and NP-complete for SEFS.

2 Preliminaries

For a finite set Δ, $|\Delta|$ denotes the cardinality of Δ. Let Σ be a finite alphabet and X be a set of variables. We assume that Σ and X are mutually disjoint. We denote by Σ^* the set of all words over Σ and by ε the empty word. A *pattern* is an element of $(\Sigma \cup X)^*$. For a pattern π, $var(\pi)$ denote the set of all variables in π. The *length* of a pattern π is denoted by $|\pi|$. Let Π be a finite alphabet of *predicate symbols* associated with a mapping $r : \Pi \to \mathbb{N}$, called *arity*. An *atom* is an expression of the form $p(\tau_1, \ldots, \tau_{r(p)})$, where $p \in \Pi$ and $\tau_1, \ldots, \tau_{r(p)}$ are patterns. For an atom $A = p(\pi_1, \ldots, \pi_n)$, the length of A is defined by $|A| = |\pi_1| + \cdots + |\pi_n|$. An *axiom* C is an expression of the form $A \leftarrow B_1, \ldots, B_m$, where $m \geq 0$ and $A, B_1, \ldots, B_m$ are atoms. The parts A and $B_1, \ldots, B_m$ are called the *head* and the *body* of C, respectively. In case $m = 0$, the axiom C is called a *unit* axiom. We write A for a unit axiom $A \leftarrow$.

An *elementary formal system* (EFS) is a quadruple $S = (\Sigma, \Pi, \Gamma, p_0)$, where Γ is a finite set of axioms and $p_0 \in \Pi$ is the distinguished predicate symbol. For convention, we often identify Γ with $(\Sigma, \Pi, \Gamma, p_0)$ if Σ, Π, and p_0 are understood from context.

Example 1. The quadruple $S = (\{a, b\}, \{p_0\}, \Gamma, p_0)$ is an EFS, where

$$\Gamma = \left\{ \begin{array}{ll} p_0(xy) & \leftarrow p_0(x), p_0(y), \\ p_0(axb) & \leftarrow p_0(x), \\ p_0(\varepsilon). & \end{array} \right\}$$

A *substitution* θ is a homomorphism $\theta : (\Sigma \cup X)^* \to (\Sigma \cup X)^*$ such that $\theta(a) = a$ for any $a \in \Sigma$. The substitution that maps x_1 to t_1, ..., x_m to t_m is denoted by $\{x_1 := t_1, \ldots, x_m := t_m\}$. A substitution θ is *erasing* if $\theta(x) = \varepsilon$ for some variable x, and *nonerasing* otherwise. In this paper, erasing substitutions are allowed unlike [1]. For a pattern π and a substitution θ, we denote by $\pi\theta$ the image of π with θ. For an atom $A = p(\pi_1, \ldots, \pi_n)$ and an axiom $C = A \leftarrow B_1, \ldots, B_m$, we extend θ by defining $A\theta = p(\pi_1\theta, \ldots, \pi_n\theta)$ and $C\theta = A\theta \leftarrow B_1\theta, \ldots, B_m\theta$.

Definition 1. Let $S = (\Sigma, \Pi, \Gamma, p_0)$ be an EFS. We define a binary relation $\vdash$ inductively as follows:

(i) If C is an axiom in Γ then $\Gamma \vdash C$.

(ii) If $\Gamma \vdash C$ then $\Gamma \vdash C\theta$ for any substitution θ.

(iii) If $\Gamma \vdash A \leftarrow B_1, \ldots, B_m, B_{m+1}$ and $\Gamma \vdash B_{m+1}$ then $\Gamma \vdash A \leftarrow B_1, \ldots, B_m$.

If $\Gamma \vdash C$ then we say C is *provable* from Γ. We define $L(S) = \{w \in \Sigma^* \mid \Gamma \vdash p_0(w)\}$. A language $L \subseteq \Sigma^*$ is *definable* by EFS or an *EFS language* if such S exists.

Example 2. The Dyck language over $\{a, b\}$ is the language L_1 defined by the context-free grammar with the productions $\{S \Rightarrow SS, S \Rightarrow aSb, S \Rightarrow \varepsilon\}$. L_1 is definable by the EFS of Example 1. We can see $aaabbabb \in L_1$ by the *proof* in Figure 1.

Now, we introduce some constraints on the patterns of EFS's. An axiom

$$q(\pi_1, \ldots, \pi_n) \leftarrow q_1(\tau_1, \ldots, \tau_{t_1}), q_2(\tau_{t_1+1}, \ldots, \tau_{t_2}), \ldots, q_l(\tau_{t_{l-1}+1}, \ldots, \tau_{t_l})$$

is *hereditary* if, for each $1 \leq j \leq t_l$, a pattern τ_j is a subword of some π_i [9, 10]. By definition, a unit axiom is hereditary. A hereditary axiom $A \leftarrow B_1, \ldots, B_m$ is *length-bounded* if $|A\theta| \geq |B_1\theta| + \cdots + |B_m\theta|$ for any substitution θ [3, 11]. A hereditary axiom is *simple* if it is of the form $p(\pi) \leftarrow q_1(x_1), \ldots, q_m(x_m)$, where π is a pattern and $x_1, \ldots, x_m$ are mutually distinct variables in π [2, 3].

$$
\begin{array}{ll}
C_1 = p_0(\varepsilon) & \text{(axiom)} \\
C_2 = p_0(axb) \leftarrow p_0(x) & \text{(axiom)} \\
C_3 = p_0(ab) \leftarrow p_0(\varepsilon) & (C_2\{x := \varepsilon\}) \\
C_4 = p_0(ab) & (C_1, C_3) \\
C_5 = p_0(aabb) \leftarrow p_0(ab) & (C_2\{x := ab\}) \\
C_6 = p_0(aabb) & (C_4, C_5) \\
C_7 = p_0(xy) \leftarrow p_0(x), p_0(y) & \text{(axiom)} \\
C_8 = p_0(aabbab) \leftarrow p_0(aabb), p_0(ab) & (C_7\{x := aabb, y := ab\}) \\
C_9 = p_0(aabbab) & (C_4, C_6, C_8) \\
C_{10} = p_0(aaabbabb) \leftarrow p_0(aabbab) & (C_2\{x := aabbab\}) \\
C_{11} = p_0(aaabbabb) & (C_9, C_{10})
\end{array}
$$

Figure 1: The proof for $\Gamma_1 \vdash p_0(aaabbabb)$

An EFS is *hereditary* (resp. *length-bounded* and *simple*) if all axioms are hereditary (resp. length-bounded and simple). We denote by HEFS (resp. LB-HEFS and SEFS) the class of hereditary EFS's (resp. length-bounded EFS's and simple EFS's) and the class of corresponding languages.

By definition, SEFS and LB-HEFS are subclasses of HEFS. From Arikawa *et al.* [3], we have the hierarchy CFL $\subseteq$ SEFS $\subseteq$ LB-HEFS $\subseteq$ CSL, where CFL and CSL are the classes of context-free languages and context-sensitive languages, respectively.

Example 3. The language $\{a^n b^n c^n \mid n \geq 1\}$ is definable by the HEFS Γ_2 with the distinguished monadic predicate p_0:

$$
\Gamma_2 = \left\{ \begin{array}{ll} p_0(xyz) & \leftarrow p(x, y, z), \\ p(ax, by, cz) & \leftarrow p(x, y, z), \\ p(a, b, c). & \end{array} \right\}
$$

3 The simulating capacity of SEFS's

In this section, we demonstrate that SEFS's can simulate languages generated by various grammatical devices based on patterns. Thus, SEFS's provide a uniform framework to study these devices.

First, we examine *pattern languages* (PL) [1] and *multi-pattern languages* (MPL) which are extensions of pattern languages to unions [6]. For a pattern π and an alphabet Σ, the pattern language $L_{E,\Sigma}(\pi)$ is the set $\{\pi\theta \in \Sigma^* \mid \theta \text{ is any substitution}\}$. The language is definable by the SEFS $\Gamma = \{p_0(\pi)\}$. For a set of patterns $\{\pi_1, \ldots, \pi_n\}$ and an alphabet Σ, a multi-pattern language $L_{E,\Sigma}(\pi_1, \ldots, \pi_n)$ is the set $\cup_{i=1}^{n} L_{E,\Sigma}(\pi_i)$. The language is also definable by the SEFS $\Gamma = \{p_0(\pi_i) \mid 1 \leq i \leq n\}$.

Next, we consider grammars that produce strings by iteratively substituting strings for a set of patterns in a non-synchronous way. A *pattern grammar* (GPL) [5] is a pair $G = (P, A)$ of a finite set P of patterns and a finite set $A \subseteq \Sigma^*$ of strings. The language defined by G is $L(G) = \bigcup_{i \geq 0} D_i$, where $D_0 = A$ and

$$D_{i+1} = \bigcup_{\pi \in P} \{\pi\theta \in \Sigma^* \mid x\theta \in D_0 \cup \cdots \cup D_i \text{ for all } x \in var(\pi)\}.$$

A *pattern system* [8] is a quadruple $G = (\Sigma, V, p, t)$, where $n \geq 0$, $V = \{X_1, \ldots, X_n\}$ is a set of variables, and p and t are mappings from V into nonempty finite sets of patterns in $(\Sigma \cup V)^* V (\Sigma \cup V)^*$ and strings in Σ^*, respectively.

Definition 2. (Mitrana *et al.* [8]) Let $G = (\Sigma, V, p, t)$ be a pattern system and $n = |V|$. Then, for some fixed j, a *non-synchronized pattern language* (NSPL) is the language defined by $L_{NS}(G, j) = \cup_{i \geq 0} D_j^{(i)}(G)$, where $D_j^{(i)}(G)$ is defined recursively as follows: $D_j^{(0)}(G) = t(X_j)$, and

$$D_j^{(i+1)}(G) = \bigcup_{k \leq i} D_j^{(k)}(G) \cup \{\pi_j\theta \mid \pi_j \in p(X_j), X_l\theta \in D_l^{(i)}(G), 1 \leq l \leq n\}.$$

We denote by PL, MPL, GPL and NSPL the corresponding classes of languages defined by the grammatical devices introduced above. From Mitrana *et al.* [8], we know that two inclusions PL $\subset$ MPL $\subset$ NSPL and GPL $\subset$ NSPL hold.

The only relation among SEFS and Chomsky hierarchy known before is the inclusion CFL $\subseteq$ SEFS $\subseteq$ CSL. Theorem 1 precisely locates SEFS in the hierarchy consisting of grammatical devices based on patterns.

Lemma 1. Any language definable by SEFS is an NSPL.

Proof: Let $C = p(\pi) \leftarrow p_1(x_1), \ldots, p_m(x_m)$ be a simple axiom. If there exist variables $x_{m+1}, \ldots, x_{m+n}$ of π not appearing in the body then we consider the set of axioms $\Gamma_0 = \{r(a) \mid a \in \Sigma\} \cup \{r(ax) \leftarrow r(x) \mid a \in \Sigma\}$ and the axiom

$$C' = p(\pi) \leftarrow p_1(x_1), \ldots, p_m(x_m), r(x_{m+1}), \ldots, r(x_{m+n}).$$

It is easy to see $\Sigma^* = \{w \in \Sigma^* \mid \Gamma_0 \vdash r(w)\}$. Thus all axioms in $\Gamma_0 \cup \{C'\}$ are equivalent to the axiom C. Therefore we can assume that, for any simple axiom, all variables in the head appear in the body.

Let $S = (\Sigma, \Pi, \Gamma, p_0)$ be an SEFS. First, we shall transform S into an equivalent SEFS $S' = (\Sigma, \Pi', \Gamma', q_0)$ in which all predicate symbols in the body of any axiom are different from each other. Let n be the maximum of $|var(\pi)|$ for all pattern π appearing in Γ. Let $\Pi' = \{q^{(i)} \mid q \in \Pi, 1 \leq i \leq n\} \cup \{q_0\}$, where $q^{(i)}$ is the new predicate symbol

superscripted with the index i. For each unit axiom $p(w) \in \Gamma$ and $1 \leq i \leq n$, the set Γ' contains $p^{(i)}(w)$. For each $p(\pi) \leftarrow p_1(x_1), \ldots, p_m(x_m) \in \Gamma$, Γ' contains the axioms

$$p^{(i)}(\pi) \leftarrow p_1^{(1)}(x_1), \ldots, p_m^{(m)}(x_m),$$

for each $1 \leq i \leq n$. Finally we add axioms $q_0(x) \leftarrow p_0^{(1)}(x), \ldots, p_0^{(n)}(x)$ into Γ'. By induction on the construction of proofs, we can see that $\Gamma \vdash p(w)$ if and only if $\Gamma' \vdash p^{(i)}(w)$ for any $p \in \Pi$, $w \in \Sigma^*$, and any $1 \leq i \leq n$. Thus, we have $L(S) = L(S')$.

Next, we built a pattern system $G = (\Sigma, V, p, t)$ as follows. V is the set $\{X_q \mid q \in \Pi'\}$ of variables. For each unit axiom $q(w) \in \Gamma'$, the set $t(X_p)$ contains w. Note that w contains no variables because any variable in the head must appear in the body. For each non-unit axiom $q(\pi) \leftarrow q_1(x_1), \ldots, q_m(x_m) \in \Gamma'$, the set $p(X_q)$ contains the pattern $\pi' = \pi\{x_1 := X_{q_1}, \ldots, x_m := X_{q_m}\} \in (\Sigma \cup V)^*$. Note that if $q_1, \ldots, q_m$ are mutually distinct then π' is a renaming variant of π. Therefore, it is not difficult to see that $\{w \in \Sigma^* \mid \Gamma' \vdash q(w)\} = L_{NS}(G, q)$ for each index $q \in \Pi'$. Combining these results, we conclude that the pattern system G defines $L(S')$ with the variable X_{q_0}. Hence, it immediately follows that $L(S) = L_{NS}(G, q_0)$. □

Theorem 1. NSPL is precisely the class of languages definable by SEFS.

Proof: By Lemma 1, it remains to show that any NSPL is definable by SEFS. Suppose L is defined by a pattern system $G = (\Sigma, V, p, t)$ with variable X_i, where $V = \{X_1, \ldots X_n\}$. We shall built an SEFS $S = (\Sigma, \Pi, \Gamma, p_i)$ such that $L(S) = L_{NS}(G, i)$ as follows. Π consists of n predicate symbols $p_1, \ldots, p_n$ that correspond to the variables $X_1, \ldots, X_n$ in V. For each $1 \leq j \leq n$ and each $w \in t(X_j)$, Γ contains $p_j(w)$. For each $1 \leq j \leq n$ and each $\pi \in p(X_j)$, Γ contains $p_j(\pi') \leftarrow p_{j_1}(x_1), \ldots, p_{j_m}(x_m)$, where $var(\pi) = \{X_{j_1}, \ldots, X_{j_m}\}$ and $\pi' = \pi\{X_{j_1} := x_1, \ldots, X_{j_m} := x_m\}$. It is straightforward to show that $L_{NS}(G, i) = L(S)$ for arbitrary index i. □

Mitrana *et al.* raised a question whether NSPL $\subseteq$ EOL holds [8]. Concerning with EOL, Sudborough proved that EOL is LOGCFL-complete [13]. On the other hand, we know that SEFS = NSPL is also LOGCFL-complete from the inclusion CFL $\subseteq$ SEFS $\subseteq$ LOGCFL shown by Miyano *et al.* [9]. From these observations and Theorem 1 above, we know both NSPL and EOL belong to LOGCFL, and they are complete for the class. Thus, we know that NSPL and EOL are somewhat similar in computational complexity.

Corollary 2. NSPL and EOL are equivalent under many-one logspace reduction.

The following theorem is straightforward from the proof of Theorem 1.

Theorem 3. GPL is precisely the class of languages definable by SEFS whose predicate symbol is only p_0.

4 The expressive power of HEFS's

In the previous section, we considered restricted HEFS's, called SEFS's. SEFS's are less powerful than HEFS's since $L = \{a^n b^n c^n \mid n \geq 1\} \in$ HEFS but $L \notin$ SEFS $=$ NSPL [8]. In this section, we examine the expressive power of non-restricted HEFS's and show that HEFS is exactly the class of languages accepted by deterministic Turing machines in polynomial time.

Miyano *et al.* showed the following lemma using the property that the number of subwords of any input is bounded by some polynomial in the length of the input.

Lemma 2. (Miyano *et al.* [10]) Any language definable by HEFS is accepted by some deterministic Turing machine in polynomial time.

To show the converse of Lemma 2, we use that any language in P is accepted by some two-way alternating finite automaton with k heads (2AFA(k)) [7]. A 2AFA(k) M is a finite automaton which has the single input tape, two-way access to the tape, and k read-only heads. An input on the tape is enclosed with the left endmarker ¢ and right endmarker \$. A state of M is either *existential* or *universal*. A *configuration* of M on input $w \in \Sigma^*$ is a $(k+1)$-tuple $(q, h_1, \ldots, h_k)$, where q is a state of M and h_j is the position of the jth head $(0 \leq h_j \leq |w| + 1)$ for each $1 \leq j \leq k$. The configuration is existential (resp. universal) if q is existential (resp. universal).

Let K be the set of states of M. A *transition function* is a mapping from $K \times (\Sigma \cup \{¢, \$\})^k$ into the subsets of $K \times \{L, N, R\})^k$. Let C be a configuration of M and $a_j \in \Sigma \cup \{¢, \$\}$ be the jth symbol of the tape of C. Then a *transition* from C is defined by $\delta(p, a_1, \ldots, a_k) \ni (q, d_1, \ldots, d_k)$, which means that M changes the state into q and moves the head toward the direction $d_j \in \{L, N, R\}$, where L, N and R stand for left, neutral, and right, respectively. A configuration followed from C is called a *successor* of C. Let $D_1, \ldots, D_m$ be all successors of C. When C is existential (resp. universal), C leads to acceptance if and only if D_i leads to acceptance some (resp. all) i.

For a configuration C on input $w \in \Sigma^*$, we define an atom conf(C) as follows. The $(2k+1)$-ary predicate symbol of conf(C) is p subscripted with $e \in \{0,1\}^k$. We denote by $e[j]$ the jth bit of e. In the rest of this section, we write $p_e(\pi_1, \ldots, \pi_{2k}; \pi_{2k+1})$ for the atom $p_e(\pi_1, \ldots, \pi_{2k}, \pi_{2k+1})$. Then conf($C$) is defined by

$$\mathrm{conf}(C) = p_e(u_1, v_1, \ldots, u_k, v_k; w),$$

where for each $1 \leq j \leq k$, u_j and v_j are subwords of w such that $u_j v_j = w$, if $h_j \geq 1$ then $|u_j| = h_j - 1$ and $e[j] = 1$, and if $h_j = 0$ then $|u_j| = 0$ and $e[j] = 0$. Each triplet $(e[j], u_j, v_j)$ denotes the position of the jth head. For example, we represent the configuration $(p, 3, 0, 1)$ of 2AFA(3) on input *abba* as the atom $p_{101}(ab, ba, \varepsilon, abba, \varepsilon, abba; abba)$.

Now we prove the converse of Lemma 2.

Theorem 4. Any language $L \in \mathrm{P}$ is definable by HEFS.

Proof: Let M be a 2AFA(k) that accepts L. We construct an EFS $S = (\Sigma, \Pi_M, \Gamma_M, p_0)$ as follows. First, the set Π_M is defined as

$$\Pi_M = \{p_0\} \cup \{p_e \mid p \text{ is a state of } M \text{ and } e \in \{0,1\}^k\},$$

where $r(p_0) = 1$ and $r(p_e) = 2k+1$.

Next, we construct the set Γ_M of axioms from transitions of M. Let $(q, d_1, \ldots, d_k) \in \delta(p, a_1, \ldots, a_k)$ be a transition of M and p be an existential state. We can assume that M moves at most one head, say the sth head. Let E be the set of all pairs $(e, e') \in \{0,1\}^k \times \{0,1\}^k$ such that $e[j] = e'[j]$ for all $j \neq s$ and if $d_s = R$ (resp. $d_s = L$) then $e'[s] = 1$ (resp. $e[s] = 1$). The set E corresponds to the restriction that any head of M does not go out of the tape. For all $(e, e') \in E$, we add the axioms

$$p_e(\pi_1, \tau_1, \ldots, \pi_k, \tau_k; \pi_s\tau_s) \leftarrow q_{e'}(\pi'_1, \tau'_1, \ldots, \pi'_k, \tau'_k; \pi_s\tau_s), \tag{1}$$

into Γ_M, where each (π_j, τ_j) and (π'_j, τ'_j) are defined as follows: If the sth head moves right ($d_s = R$) then

$$(\pi_s, \tau_s, \pi'_s, \tau'_s) = \begin{cases} (x_s, a_s y_s, x_s a_s, y_s) & \text{if } e[s] = 1, \\ (\varepsilon, y_s, \varepsilon, y_s) & \text{if } e[s] = 0. \end{cases}$$

If the sth head moves left ($d_s = L$) then

$$(\pi_s, \tau_s, \pi'_s, \tau'_s) = \begin{cases} (x_s b, a_s y_s, x_s, b a_s y_s) & \text{if } e'[s] = 1, a_s \neq \$, \\ (x_s b, \varepsilon, x_s, b) & \text{if } e'[s] = 1, a_s = \$, \\ (\varepsilon, a_s y_s, \varepsilon, a_s y_s) & \text{if } e'[s] = 0. \end{cases}$$

For all $1 \leq j \leq k$ such that $d_j = N$,

$$(\pi_j, \tau_j, \pi'_j, \tau'_j) = \begin{cases} (x_j, a_j y_j, x_j, a_j y_j) & \text{if } e[j] = 1, a_j \neq \$, \\ (x_j, \varepsilon, x_j, \varepsilon) & \text{if } e[j] = 1, a_j = \$, \\ (\varepsilon, y_j, \varepsilon, y_j) & \text{if } e[j] = 0. \end{cases}$$

Let p be a universal state. We can assume that M does not move its heads but change its state p into some states $q^{(1)}, \ldots, q^{(m)}$ universally. This transition is translated into the following axioms for all $e \in \{0,1\}^k$:

$$\begin{aligned} &p_e(\pi_1, \tau_1, \ldots, \pi_k, \tau_k; z) \leftarrow \\ &\quad q_e^{(1)}(\pi_1, \tau_1, \ldots, \pi_k, \tau_k; z), \ldots, q_e^{(m)}(\pi_1, \tau_1, \ldots, \pi_k, \tau_k; z), \end{aligned} \tag{2}$$

where z is a variable and for all $1 \leq j \leq k$,

$$(\pi_j, \tau_j, \pi_j, \tau_j) = \begin{cases} (x_j, a_j y_j, x_j, a_j y_j) & \text{if } e[j] = 1, a_j \neq \$, \\ (x_j, \varepsilon, x_j, \varepsilon) & \text{if } e[j] = 1, a_j = \$, \\ (\varepsilon, y_j, \varepsilon, y_j) & \text{if } e[j] = 0. \end{cases}$$

Finally, for the initial state q of M, we add the axiom with $e = 1^k$

$$p_0(x) \leftarrow q_e(\varepsilon, x, \ldots, \varepsilon, x; x), \tag{3}$$

and for all accepting states p and for all $e \in \{0,1\}^k$, we add axioms

$$p_e(x_1, y_1, \ldots, x_k, y_k; z). \tag{4}$$

The axioms of (1) are hereditary since $\pi'_j = \pi_j, \tau'_j = \tau_j$ for all $j \neq s$, and π_s, τ_s are subwords of $\pi_s\tau_s$ which is the last argument in the head. Those of (2), (3) and (4) are also hereditary since ε is a subword of any string and an axiom without the body is hereditary. Thus S is hereditary.

Claim: A configuration $C = (p, h_1, \ldots, h_k)$ of M on input $w \in \Sigma^*$ leads to acceptance if and only if $\Gamma_M \vdash \mathrm{conf}(C)$.

Proof of Claim: Assume that C leads to acceptance. If C is an accepting configuration then $\Gamma_M \vdash \mathrm{conf}(C)$ from (4).

If C is existential then there exists a successor D of C such that D leads to acceptance. It is sufficient to show $\mathrm{conf}(C) \leftarrow \mathrm{conf}(D)$ is an instance of some axiom in Γ_M since $\Gamma_M \vdash \mathrm{conf}(D)$ by the induction hypothesis. There exists an axiom $A \leftarrow B \in \Gamma_M$ such that $\mathrm{conf}(C) = A\theta$ for some θ. Therefore we show $B\theta = \mathrm{conf}(D)$. Let $(q, d_1, \ldots, d_k) \in \delta(p, a_1, \ldots, a_k)$ be a transition that changes C into D. From the construction of (1), $(\pi_j, \tau_j) = (\pi'_j, \tau'_j)$ for all j such that $d_j = N$. Thus, these pairs simulate the move of the jth head. If $h_s = 0$ and $d_s = R$ then $(\pi'_s, \tau'_s) = (\pi_s, \tau_s) = (\varepsilon, y_s)$ and $(e[s], e'[s]) = (0, 1)$ from (1). Thus the pairs simulate the move of the sth head. Otherwise it is not difficult to show that these pairs simulate the move of the sth head. Thus, $B\theta = \mathrm{conf}(D)$.

If C is universal and has the successors $C_1, \ldots, C_m$ then all C_i lead to acceptance. By the induction hypothesis, $\Gamma_M \vdash \mathrm{conf}(C_i)$ for all i and there exists $\mathrm{conf}(C) \leftarrow \mathrm{conf}(C_1), \ldots, \mathrm{conf}(C_m)$ as an instance of some axiom in Γ_M. Thus, we have $\Gamma_M \vdash \mathrm{conf}(C)$.

The converse direction is proved in a similar way by using induction on the construction of the proof for $\Gamma_M \vdash \mathrm{conf}(C)$.(*End of the Proof of Claim*)

From the above claim, the initial configuration C_0 leads to acceptance if and only if $\Gamma_M \vdash \mathrm{conf}(C_0)$. Hence, $w \in L$ if and only if $w \in L(S)$ because Γ_M contains the axiom $p_0(x) \leftarrow q_e(\varepsilon, x, \ldots, \varepsilon, x; x)$ and $\mathrm{conf}(C_0) = q_e(\varepsilon, w, \ldots, \varepsilon, w; w)$ where $e = 1^k$ and q is the initial state of M. This completes the proof. □

From Lemma 2 and Theorem 4, we have the main theorem.

Theorem 5. P is exactly the class of languages definable by HEFS.

A *linear HEFS* is an HEFS such that each axiom has at most one atom in the body. By a similar proof for Theorem 4, we have the following corollary.

Corollary 6. NSPACE($\log n$) is exactly the class of languages definable by linear HEFS.

5 Complexity of the membership problem

In this section, we investigate the complexity of the membership problems for HEFS and SEFS. The *membership problem* for a class C of grammars is the problem of, given a string $w \in \Sigma^*$ and a grammar $G \in C$, deciding whether $w \in L(G)$. Let EXPTIME be the class of languages accepted by deterministic Turing machines in time $O(2^{n^c})$ for some $c > 0$.

Theorem 7. The membership problem for HEFS is EXPTIME-complete.

Proof: For one direction, we will give an *alternating Turing machine* (ATM) M that solves the membership problem in space $O(n^c)$ for some $c > 0$. This can be done in space $O(r \log n)$ by using pointers on input to represent an atom and by using alternations to simulate a top down proof for $\Gamma \vdash p_0(w)$ based on techniques in [10], where r is the maximum arity of predicates. Since EXPTIME is precisely the class of languages accepted by ATM's in polynomial space [4], we know that the membership problem is in EXPTIME.

For the converse direction, let $L \subseteq \Sigma^*$ be a language in EXPTIME. Then we can assume that there is an ATM $M = (\{0,1\}, \Delta, Q, \delta, q_0, F, B, U)$ with only one work/input tape that accepts L in space $O(n^c)$ for some $c > 0$ [4]. An idea is to encode a configuration $(a_1 \cdots a_{i-1} p\, a_i \cdots a_n)$ of M, where M is scanning the ith cell in state p and $a_i \in \{0,1,B\}$ for all $1 \leq i \leq n$, by an atom $p(a_1, \ldots, a_{i-1}, \uparrow, a_i, \ldots, a_n, 0, 1, B)$ of arity $n^c + 4$ over the alphabet $\{0, 1, B, \uparrow\}$.

Now, we built an HEFS $S = (\{0,1,B,\uparrow\}, \Pi = Q, \Gamma, p_0)$. For convention, we denote by $x^{(i)}$ the sequence $(x_1, \ldots, x_i)$ of mutually distinct variables. An existential transition $(q, b, d) \in \delta(p, a)$ is represented by the axiom in Γ

$$\begin{array}{ll} p(x^{(i-1)}, \uparrow, a, y^{(n^c-i)}, 0, 1, B) \leftarrow q(x^{(i-1)}, b, \uparrow, y^{(n^c-i)}, 0, 1, B) & \text{if } d = R, \\ p(x^{(i-2)}, x_{i-1} \uparrow, a, y^{(n^c-i)}, 0, 1, B) \leftarrow q(x^{(i-2)}, \uparrow, x_{i-1}, a, y^{(n^c-i)}, 0, 1, B) & \text{if } d = L, \end{array}$$

and similar constructions are possible for $d = N$. A universal transition is assumed to have the form $\{(q_1, a, 0), \ldots, (q_m, a, 0)\} = \delta(p, a)$, and is represented by the axiom

$$\begin{array}{l} p(x^{(i-1)}, \uparrow, a, y^{(n^c-i)}, 0, 1, B) \leftarrow \\ \quad p(x^{(i-1)}, \uparrow, a, y^{(n^c-i)}, 0, 1, B), \ldots, p(x^{(i-1)}, \uparrow, a, y^{(n^c-i)}, 0, 1, B). \end{array}$$

For the initial configuration, we have

$$p_0(\uparrow x_1 \cdots x_{n^c} 01B) \leftarrow p_0(\uparrow, x_1, \ldots, x_{n^c}, 0, 1, B).$$

Then, we can show that $p_0(\uparrow \$a_1 \cdots a_n B \cdots B01B) \in L(S)$ if and only if M accepts input $w = a_1 \cdots a_n$. It is straightforward to verify that this transformation from L to

the membership problem can be done in space $O(\log n)$. Combining these results, we prove the theorem. □

As a corollary, we can easily show that the membership problem for linear HEFS is PSPACE-complete. For every $k \geq 1$, let LB-HEFS(k) be the subclass of LB-HEFS for which the arity of predicates are bounded by k.

Theorem 8. For every $k \geq 1$, the membership problems for LB-HEFS(k) and SEFS are both NP-complete.

Proof: Angluin [1] showed that the membership problem for PL is NP-complete. Since any PL is definable by an LB-HEFS(1) as seen in Section 3, the membership problem for PL is reducible to that for LB-HEFS(k) for every $k \geq 1$. This is also the case for SEFS.

Let $w \in \Sigma^*$ be any string. Since an LB-HEFS Γ is hereditary and the arities of predicates are bounded by constant k, there are at most polynomially many distinct atoms whose arguments are subwords of w. Therefore, if $\Gamma \vdash p(w)$ then there is a proof T of size polynomial in the total size of w and Γ. Thus, a nondeterministic Turing machine M can decide whether $\Gamma \vdash p(w)$ by first guessing T and then checking T in polynomial time. Hence, we conclude that LB-HEFS(k) languages are decidable in NP. Since SEFS $\subseteq$ LB-HEFS(1), the result immediately follows. □

The membership problem for CFL (CFG as representation) is known to be P-complete. Hence, it is interesting that the complexity of the membership problems for CFL and SEFS are quite different in Theorem 8 above, while the languages of CFL and SEFS have almost same complexity in Corollary 2 of Section 3.

6 Conclusion

In this paper, we studied the HEFS languages. We show that HEFS captures the complexity class P and linear HEFS captures NSPACE($\log n$). We also show that SEFS and the subclass of SEFS with only one predicate symbol precisely generate NSPL and GPL, respectively. Finally, we investigate the complexity of the membership problems for HEFS and for SEFS.

For a pattern system, Mitrana *et al.* defined two languages, one is NSPL and the other is a *strongly synchronized pattern languages* (SSPL) [8]. Both languages are generated by iterative substitutions. At some step of iterations, all substituted words are generated in the previous step for SSPL, while they are generated in before steps for NSPL. The classes SSPL and NSPL are incomparable [8]. Thus, it is an interesting task to find a subclass of HEFS that corresponds to SSPL. Since SSPL is closely related with some OL system, it is also interesting to compare HEFS with them.

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