Prediction-Preserving Reducibility with Membership Queries on Formal Languages^{*}

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Abstract. This paper presents the *prediction-preserving reducibility with membership queries* (*pwm-reducibility*) on formal languages, in particular, simple CFGs and finite unions of regular pattern languages. For the former, we mainly show that DNF formulas are pwm-reducible to CFGs that is sequential or that contains at most one nonterminal. For the latter, we show that both bounded finite unions of *regular* pattern languages and unbounded finite unions of *substring* pattern languages are pwm-reducible to DFAs.

Keywords: prediction-preserving reduction with membership queries, prediction with membership queries, context-free grammars, pattern languages, grammatical inference, learning theory.

1 Introduction

The task of predicting the classification of a new example is frequently discussed from the viewpoints of both *passive* and *active* settings. In a passive setting, the examples are all chosen independently according to a fixed but unknown probability distribution, and the learner has no control over selection of examples [12, 18]. In an active setting, on the other hand, the learner is allowed to ask about particular examples, that is, the learner makes *membership queries*, before the new example to predict is given to the learner [3, 6].

Concerned with language learning, we can design a polynomial-time algorithm to predict deterministic finite automata (DFAs) in an active setting [3],

^{*} This work is partially supported by Grant-in-Aid for Scientific Research (B) nos. 11558040, 11558084 and 13558036, and Grant-in-Aid for Young Scientists (A) no. 12780233. Also this paper will be published in Proc. of 13th International Symposium on Fundamentals of Computation Theory (FCT2001), Lecture Notes in Computer Science ©Springer-Verlag.

while predicting DFAs is as hard as computing certain apparently hard cryptographic predicates in a passive setting [12]. Furthermore, predicting nondeterministic finite automata (NFAs) and context-free grammars (CFGs) is also hard under the same cryptographic assumptions in an active setting [6]. Here, the *cryptographic assumptions* denote the intractability of inverting RSA encryption, recognizing quadratic residues and factoring Blum integers.

Pitt and Warmuth [18] have been formalized the model of a prediction and a reduction between two prediction problems that preserves polynomial-time predictability called the *prediction-preserving reduction* in a passive setting. Angluin and Kharitonov [6] have extended to the prediction and the reduction in an active setting. The reduction is called the *prediction-preserving reduction with membership queries* or a *pwm-reduction* for short.

The prediction is an weaker learning model than PAC-learning or query learning models [5, 6, 18]; If a class is polynomial-time learnable with equivalence (and membership) queries, then it is polynomial-time PAC-learnable (with membership queries), and if a class is polynomial-time PAC-learnable (with membership queries), then it is polynomial-time predictable (with membership queries).

Except the above general results, the detailed results of the pwm-reducibility on formal languages, for example, the restricted CFGs or another languages such as pattern languages, have few found elsewhere. Furthermore, many researchers have been interested in the pwm-reducibility on Boolean concepts but not on formal languages [6, 12, 18]. Hence, in this paper, we pay our attention to the pwm-reducibility on simple CFGs and finite unions of regular pattern languages.

For the former classes, we introduce several simple CFGs as follows: the *linear grammars* ($\mathcal{L}_{\text{linear}}$), the *right-linear grammars* ($\mathcal{L}_{\text{right-linear}}$), and the *left-linear grammars* ($\mathcal{L}_{\text{left-linear}}$) as usual; the *k-bounded CFGs* [4] ($\mathcal{L}_{k\text{-bounded-CFG}}$) each of which right-hand side of productions contains at most k nonterminals; the *sequential CFGs* [7, 24] ($\mathcal{L}_{\text{sqCFG}}$) that the set of nonterminals has a partial order $\leq$ such that $T \rightarrow vUw$ iff $T \leq U$ for nonterminals T and U ; the *properly sequential or loop-free CFGs* ($\mathcal{L}_{\text{psqCFG}}$) that is sequential but disallowing to occur the same nonterminal in both left- and right-hand sides for each production; the *k-CFGs* ($\mathcal{L}_{k\text{-CFG}}$) that contains at most k nonterminals; the *parenthesis grammars* [16] ($\mathcal{L}_{\text{paren}}$) that each production is of the form $T \rightarrow [w]$.

On the other hand, a *pattern* [2] is a string consisting of constant symbols and variables, a *regular pattern* [25] is a pattern in which each variable occurs at most once, and a *substring pattern* [26] is a regular pattern of the form xwy for a constant string w and variables x and y . The *language of a pattern* is the set of constant strings obtained by substituting nonempty constant strings for variables in the pattern. Then, for the latter classes, we deal with the *bounded finite union* of regular pattern languages by some constant m ($\mathcal{L}_{\cup_m \text{RP}}$) and the *unbounded finite union* of regular or substring pattern languages ($\mathcal{L}_{\cup \text{RP}}$ or $\mathcal{L}_{\cup \text{subP}}$).

We denote that $\mathcal{L}_1$ is pwm-reducible to $\mathcal{L}_2$ by $\mathcal{L}_1 \trianglelefteq_{\text{pwm}} \mathcal{L}_2$, and that $\mathcal{L}_1 \trianglelefteq_{\text{pwm}} \mathcal{L}_2$ and $\mathcal{L}_2 \trianglelefteq_{\text{pwm}} \mathcal{L}_1$ by $\mathcal{L}_1 \cong_{\text{pwm}} \mathcal{L}_2$. Then, this paper presents the results

described as Fig. 1. Hence, we can conclude the following polynomial-time predictability with membership queries.

$\mathcal{L}_{\text{NFA}}$	$\cong_{\text{pwm}}$	$\mathcal{L}_{\text{right-linear}}, \mathcal{L}_{\text{left-linear}}$
$\mathcal{L}_{\text{NFA}}$	$\trianglelefteq_{\text{pwm}}$	$\mathcal{L}_{\text{linear}}, \mathcal{L}_{k\text{-bounded-CFG}}$
$\mathcal{L}_{\text{DNF}}$	$\trianglelefteq_{\text{pwm}}$	$\mathcal{L}_{\text{sqCFG}}, \mathcal{L}_{\text{psqCFG}}, \mathcal{L}_{k\text{-CFG}}$
$\mathcal{L}_{\text{UDFA}}$	$\trianglelefteq_{\text{pwm}}$	$\mathcal{L}_{\text{paren}}$
		$\mathcal{L}_{\cup_m \text{RP}}, \mathcal{L}_{\cup \text{subP}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$

Fig. 1. The pwm-reducibility

1. $\mathcal{L}_{\text{linear}}, \mathcal{L}_{\text{right-linear}}, \mathcal{L}_{\text{left-linear}}, \mathcal{L}_{k\text{-bounded-CFG}}$ ($k \geq 1$) and $\mathcal{L}_{\text{paren}}$ are not polynomial-time predictable with membership queries under the cryptographic assumptions.
2. If $\mathcal{L}_{\text{sqCFG}}, \mathcal{L}_{\text{psqCFG}}$ and $\mathcal{L}_{k\text{-CFG}}$ ($k \geq 1$) are polynomial-time predictable with membership queries, then so are DNF formulas.
3. $\mathcal{L}_{\cup_m \text{RP}}$ ($m \geq 0$) and $\mathcal{L}_{\cup \text{subP}}$ are polynomial-time predictable with membership queries.

Concerned with the statement 2, this paper corrects and extends some results obtained by the paper [23]. Furthermore, concerned with the statement 3, we can show that $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\cup \text{RP}}$, so if $\mathcal{L}_{\cup \text{RP}}$ is polynomial-time predictable with membership queries, then so are DNF formulas [22].

2 Preliminaries

Let Σ and N be two non-empty finite sets of symbols such that $\Sigma \cap N = \emptyset$. A *production* $A \rightarrow \alpha$ on Σ and N is an association from a nonterminal $A \in N$ to a string $\alpha \in (N \cup \Sigma)^*$. A *context-free grammar* (*CFG*, for short) is a 4-tuple (N, Σ, P, S) , where $S \in N$ is the distinguished *start symbol* and P is a finite set of productions on Σ and N . Symbols in N are said to be *nonterminals*, while symbols in Σ *terminals*. Then, we deal with the following subclasses of CFGs.

- A *linear grammar* is a CFG $G = (N, \Sigma, P, S)$ such that each production in P is of the forms $T \rightarrow wUv$ or $T \rightarrow w$ for $T, U \in N$ and $w, v \in \Sigma^*$. In particular, a *right-linear* (*resp.*, *left-linear*) *grammar* if it is a linear grammar such that each production is of the forms either $T \rightarrow wU$ (*resp.*, $T \rightarrow Uw$) or $T \rightarrow w$ for $T, U \in N$ and $w \in \Sigma^*$.
- A CFG $G = (N, \Sigma, P, S)$ is called *k-bounded* [4] if the right-hand side of each production in P has at most k nonterminals.
- A CFG $G = (N, \Sigma, P, S)$ is called *sequential* [7, 24] if the nonterminals in N are labeled $S = T_1, \dots, T_n$ such that, for each production $T_i \rightarrow w$, $w \in (\Sigma \cup \{T_j \mid i \leq j \leq n\})^*$. In particular, a sequential CFG satisfying that

- $w \in (\Sigma \cup \{T_j \mid i < j \leq n\})^*$ for each production $T_i \rightarrow w$ is called *properly sequential* or *loop-free*.
- A CFG $G = (N, \Sigma, P, S)$ is called a *k-CFG* if $|N| \leq k$.
 - A *parenthesis grammar* [16] is a CFG $G = (N, \Sigma \cup \{[,]\}, P, S)$ such that each production in P is of the form $T \rightarrow [w]$ for $T \in N$ and $w \in (N \cup \Sigma)^*$.

Let G be a CFG (N, Σ, S, P) and α and β be strings in $(\Sigma \cup N)^*$. We denote $\alpha \Rightarrow_G \beta$ if there exist $\alpha_1, \alpha_2 \in (\Sigma \cup N)^*$ such that $\alpha = \alpha_1 X \alpha_2$, $\beta = \alpha_1 \gamma \alpha_2$ and $X \rightarrow \gamma \in P$. We extend the relation $\Rightarrow_G$ to the reflexive and transitive closure $\Rightarrow_G^*$. For a nonterminal $A \in N$, the *language* $L_G(A)$ of A is the set $\{w \in \Sigma^* \mid A \Rightarrow_G^* w\}$. The *language* $L(G)$ of G just refers to $L_G(S)$.

Next, we introduce the notions of *patterns* [2]. Let X be a countable set of *variables* such that $\Sigma \cap X = \emptyset$. A *pattern* is an element of $(\Sigma \cup X)^+$. A pattern π is called *regular* [25] if each variables in π occurs at most once. In particular, a regular pattern of the form xwy is called a *substring* pattern [26] if $x, y \in X$ and $w \in \Sigma^+$.

A *substitution* is a homomorphism from patterns to patterns that maps each symbol $a \in \Sigma$ to itself. A substitution that maps some variables to an empty string ε is called an ε -*substitution*. In this paper, we do not deal with ε -substitutions. By $\pi\theta$, we denote the image of a pattern by a substitution θ . For a pattern π , the *pattern language* $L(\pi)$ is the set $\{w \in \Sigma^+ \mid w = \pi\theta \text{ for some substitution } \theta\}$.

3 Prediction with Membership Queries

In this section, we introduce definitions and theorems for prediction and prediction-preserving reduction with membership queries due to Angluin and Kharitonov [6].

Let U denote Σ^* . If w is a string, $|w|$ denotes its length. For each $n > 0$, $U^{[n]} = \{w \in U \mid |w| \leq n\}$. A *representation of concepts* $\mathcal{L}$ is any subset of $U \times U$. We interpret an element $\langle u, w \rangle$ of $U \times U$ as consisting a *concept representation* u and an *example* w . The example w is a member of a concept u if $\langle u, w \rangle \in \mathcal{L}$. Define the *concept represented by* u as $\kappa_{\mathcal{L}}(u) = \{w \mid \langle u, w \rangle \in \mathcal{L}\}$. The *set of concepts represented by* $\mathcal{L}$ is $\{\kappa_{\mathcal{L}}(u) \mid u \in U\}$.

To represent CFGs, we define the class $\mathcal{L}_{\text{CFG}}$ as the set of pairs $\langle u, w \rangle$ such that u encodes a CFG G and $w \in L(G)$. Also we define the classes $\mathcal{L}_{\text{linear}}$, $\mathcal{L}_{\text{right-linear}}$, $\mathcal{L}_{\text{left-linear}}$, $\mathcal{L}_{k\text{-bounded-CFG}}$, $\mathcal{L}_{\text{seqCFG}}$, $\mathcal{L}_{\text{psqCFG}}$, $\mathcal{L}_{k\text{-CFG}}$, and $\mathcal{L}_{\text{paren}}$, corresponding to linear grammars, right-linear grammars, left-linear grammars, k -bounded CFGs, sequential CFGs, properly sequential CFGs, k -CFGs, and parenthesis grammars, respectively, as similar.

To represent finite unions of regular pattern languages, we define the class $\mathcal{L}_{\cup_m \text{RP}}$ as the set of pairs $\langle u, w \rangle$ such that u encodes m and a finite set $\pi_1, \dots, \pi_m$ of m regular patterns and w is in the concept represented by c iff $w \in L(\pi_i)$ for at least one π_i . Similarly, we define the class $\mathcal{L}_{\cup \text{RP}}$ (*resp.*, $\mathcal{L}_{\cup \text{subP}}$) as the set of pairs $\langle u, w \rangle$ such that u encodes a finite set $\pi_1, \dots, \pi_r$ of regular (*resp.*, substring) patterns and w is in the concept represented by c iff $w \in L(\pi_i)$ for at

least one π_i . Note that $\mathcal{L}_{\cup_m \text{RP}}$ denotes the *bounded* finite unions, whereas $\mathcal{L}_{\text{URP}}$ and $\mathcal{L}_{\text{subP}}$ denote the *unbounded* finite unions.

Additionally, we introduce the following classes. The class $\mathcal{L}_{\text{DFA}}$ (*resp.*, $\mathcal{L}_{\text{NFA}}$) denotes the set of pairs $\langle u, w \rangle$ such that u encodes a DFA (*resp.*, NFA) M and M accepts w . The class $\mathcal{L}_{\cup \text{DFA}}$ of finite union of DFAs denotes the set of pairs $\langle u, w \rangle$ such that u encodes a finite set $M_1, \dots, M_r$ of DFAs and w is in the concept represented by c iff at least one M_i accepts w . The class $\mathcal{L}_{\text{DNF}}$ denotes the set of pairs $\langle u, w \rangle$ such that u encodes a positive integer n and a DNF formula d over n Boolean variables $x_1, \dots, x_n$ such that $|w| = n$ ($w = w_1 \dots w_n$) and the assignment $x_i = w_i$ ($1 \leq i \leq n$) satisfies d .

Angluin and Kharitonov [6] have generalized the definitions of Pitt and Warmuth of prediction algorithm [18] to allow membership queries as follows.

Definition 1 (Angluin & Kharitonov [6]). A *prediction with membership queries algorithm*, or *pwm-algorithm*, is a possibly randomized algorithm A that takes as input n (a bound on the size of examples), s (a bound on the size of the target concept representations), and ϵ (an accuracy bound). It may make three different kinds of oracle calls, the responses to which are determined by the unknown target concept c_* and the unknown distribution D on $U^{[n]}$.

1. A *membership query* [3, 6] takes a string $w \in U$ as input and returns 1 if $w \in c_*$; and 0 otherwise.
2. A *request for a random classified example* takes as no input and returns a pair $\langle w, b \rangle$, where w is a string chosen independently according to D and $b = 1$ if $x \in c_*$ and $b = 0$ otherwise.
3. A *request for an element to predict* takes no input and returns a string w chosen independently according to D .

A may make any number of membership queries or requests for random classified examples, whereas A must eventually make one and only one request for an element to predict and then eventually halt with an output 0 or 1 without making any further oracle calls. The output is interpreted as A 's guess of how the target concept classifies the element returned by the request for an element to predict. A *runs in polynomial time* if its running time (counting one step per oracle call) is bounded by a polynomial in n , s and $1/\epsilon$.

Definition 2 (Angluin & Kharitonov [6]). Let $\mathcal{L}$ be a representation of concepts and c_* be the unknown target concept in $\mathcal{L}$. We say that A *successfully predicts* $\mathcal{L}$ if, for each positive integer n and s , for each positive rational ϵ , for each concept representation $u \in U^{[n]}$, for each probability distribution D on $U^{[n]}$, when A is run with input n , s and ϵ , and oracles determined by $c_* = \kappa_{\mathcal{L}}(u)$ and D , A asks membership queries that are in U and the probability in at most ϵ that the output of A is not equal to the correct classification of w by $\kappa_{\mathcal{L}}(u)$, where w is the string returned by the (unique) request for an element of predict.

Definition 3 (Angluin & Kharitonov [6]). A representation $\mathcal{L}$ of concepts is *polynomial-time predictable with membership queries* if there exists a pwm-algorithm A that runs in polynomial time and successfully predicts $\mathcal{L}$.

It is well known the following statements:

1. $\mathcal{L}_{\text{DFA}}$ is polynomial-time predictable with membership queries [3].
2. $\mathcal{L}_{\text{UDFA}}$, $\mathcal{L}_{\text{NFA}}$ and $\mathcal{L}_{\text{CFG}}$ are not polynomial-time predictable with membership queries under the cryptographic assumptions [6].
3. $\mathcal{L}_{\text{DNF}}$ is either polynomial-time predictable or not polynomial-time predictable with membership queries, if there exist one-way functions that cannot be inverted by polynomial-sized circuits [6].

Definition 4 (Angluin & Kharitonov [6]). Let $\mathcal{L}_i$ be a representation of concepts over domain U_i ($i = 1, 2$). We say that *predicting $\mathcal{L}_1$ reduces to predicting $\mathcal{L}_2$ with membership queries* (*pwm-reduces*, for short), denoted by $\mathcal{L}_1 \trianglelefteq_{\text{pwm}} \mathcal{L}_2$, if there exist an *instance mapping* $f : \mathbf{N} \times \mathbf{N} \times U_1 \rightarrow U_2$, a *concept mapping* $g : \mathbf{N} \times \mathbf{N} \times \mathcal{L}_1 \rightarrow \mathcal{L}_2$, and a *query mapping* $h : \mathbf{N} \times \mathbf{N} \times U_2 \rightarrow U_1 \cup \{\top, \perp\}$ satisfying the following conditions.

1. For each $x \in U_1^{[n]}$ and $u \in \mathcal{L}_1^{[s]}$, $x \in \kappa_{\mathcal{L}_1}(u)$ iff $f(n, s, x) \in \kappa_{\mathcal{L}_2}(g(n, s, u))$.
2. f is computable in time bounded by a polynomial in n , s and $|x|$.
3. The size of $g(n, s, u)$ is bounded by a polynomial in n , s and $|u|$.
4. For each $x' \in U_2$ and $u \in \mathcal{L}_1^{[s]}$, if $h(n, s, x') = \top$ then $x' \in \kappa_{\mathcal{L}_2}(g(n, s, u))$; if $h(n, s, x') = \perp$ then $x' \notin \kappa_{\mathcal{L}_2}(g(n, s, u))$; if $h(n, s, x') = x \in U_1$, then it holds that $x' \in \kappa_{\mathcal{L}_2}(g(n, s, u))$ iff $x \in \kappa_{\mathcal{L}_1}(u)$.
5. h is computable in time bounded by a polynomial in n , s and $|x'|$.

Furthermore, we denote that $\mathcal{L}_1 \trianglelefteq_{\text{pwm}} \mathcal{L}_2$ and $\mathcal{L}_2 \trianglelefteq_{\text{pwm}} \mathcal{L}_1$ by $\mathcal{L}_1 \cong_{\text{pwm}} \mathcal{L}_2$.

The following theorem is useful for showing the predictability or the hardness of predictability of the representations of concepts.

Theorem 1 (Angluin & Kharitonov [6]). Let $\mathcal{L}_1$ and $\mathcal{L}_2$ be representations of concepts and suppose that $\mathcal{L}_1 \trianglelefteq_{\text{pwm}} \mathcal{L}_2$.

1. If $\mathcal{L}_2$ is polynomial-time predictable with membership queries, then so is $\mathcal{L}_1$.
2. If $\mathcal{L}_1$ is not polynomial-time predictable with membership queries, then neither is $\mathcal{L}_2$.

4 Prediction-Preserving Reducibility with Membership Queries

In this section, we fix f , g and h to an instance mapping, a concept mapping, and a query mapping. Furthermore, the parameters n and s denote the bounds of examples and representations, respectively. For simplicity, we assume that the length of examples for Boolean concepts is always fixed to the upper bound n .

4.1 Simple CFGs

First of all, by using the transformation from a DFA to a right-linear grammar (cf. [8, 9]), it holds that $\mathcal{L}_{\text{DFA}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{right-linear}}$, because the size of the right-linear grammar is bounded by a polynomial in the size of a DFA. Note that the converse direction $\mathcal{L}_{\text{right-linear}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$ does not follow from the transformation from a right-linear grammar to a DFA, because the size of the DFA is not bounded by a polynomial in the size of a right-linear grammar in general.

Furthermore, we point out that $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\cup \text{DFA}}$, while $\mathcal{L}_{\text{DNF}} \not\trianglelefteq \mathcal{L}_{\text{DFA}}$ [18]. Here, the $\trianglelefteq$ means the prediction-preserving reduction without membership queries introduced by Pitt and Warmuth [18], that is, there exist f and g satisfying the requirement from 1 to 3 in Definition 4. Note that we cannot apply the same proof of $\mathcal{L}_{\text{DNF}} \trianglelefteq \mathcal{L}_{\text{DFA}}$ to proving $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$; In this case, we cannot construct h correctly.

On the other hand, by regarding the equivalent transformation between an NFA and a right-linear grammar [8, 9] as a concept mapping g , we observe that $\mathcal{L}_{\text{NFA}} \cong_{\text{pwm}} \mathcal{L}_{\text{right-linear}}$. Furthermore, for a CFG $G = (N, \Sigma, P, S)$, let G^R be a CFG (N, Σ, P', S) such that $T \rightarrow w^R \in P'$ for each $T \rightarrow w \in P$. Here, R denotes the reversal of a word. Then, for a right-linear (*resp.*, left-linear) grammar G , construct the following f , g and h :

$$\begin{aligned} f(n, s, e) &= e^R, \\ g(n, s, G) &= G^R, \\ h(n, s, e') &= e'^R. \end{aligned}$$

It is obvious that $\mathcal{L}_{\text{right-linear}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{left-linear}}$ (*resp.*, $\mathcal{L}_{\text{left-linear}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{right-linear}}$), so it holds that $\mathcal{L}_{\text{right-linear}} \cong_{\text{pwm}} \mathcal{L}_{\text{left-linear}}$. Similarly, we also observe that $\mathcal{L}_{\text{NFA}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{linear}}$ and $\mathcal{L}_{\text{NFA}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{k\text{-bounded-CFG}}$ for each $k \geq 1$. Summary:

Theorem 2. $\mathcal{L}_{\text{NFA}} \cong_{\text{pwm}} \mathcal{L}$ for $\mathcal{L} \in \{\mathcal{L}_{\text{right-linear}}, \mathcal{L}_{\text{left-linear}}\}$. Also, $\mathcal{L}_{\text{NFA}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{linear}}$ and $\mathcal{L}_{\text{NFA}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{k\text{-bounded-CFG}}$ for each $k \geq 1$.

Theorem 3. $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}$ for $\mathcal{L} \in \{\mathcal{L}_{\text{psqCFG}}, \mathcal{L}_{\text{sqCFG}}\}$.

Proof. Let d be a DNF formula $t_1 \vee \dots \vee t_m$ over n Boolean variables $x_1, \dots, x_n$. First, we define w_i^j ($1 \leq i \leq n, 1 \leq j \leq m$) as follows:

$$w_i^j = \begin{cases} 1 & \text{if } t_j \text{ contains } x_i, \\ 0 & \text{if } t_j \text{ contains } \overline{x_i}, \\ T & \text{otherwise.} \end{cases}$$

Then, construct f , g and h as follows:

$$\begin{aligned} f(n, s, e) &= e, \\ g(n, s, d) &= (\{S, T\}, \{0, 1\}, S, \{S \rightarrow w_1^1 \dots w_n^1 \mid \dots \mid w_1^m \dots w_n^m, T \rightarrow 0 \mid 1\}), \\ h(n, s, e') &= e'. \end{aligned}$$

Note that $g(n, s, d)$ is a properly sequential CFG. It is obvious that the above f , g and h satisfy the conditions of Definition 4. $\square$

Theorem 4. For each $k \geq 1$, $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{k\text{-CFG}}$.

Proof. Theorem 3 implies that $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{k\text{-CFG}}$ for each $k \geq 2$, so it is sufficient to show that $\mathcal{L}_{\text{DNF}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{1\text{-CFG}}$. Let $d = t_1 \vee \dots \vee t_m$ be a DNF formula over n Boolean variables $x_1, \dots, x_n$. First, define w_i^j ($1 \leq i \leq n, 1 \leq j \leq m$) as follows:

$$w_i^j = \begin{cases} 1 & \text{if } t_j \text{ contains } x_i, \\ 0 & \text{if } t_j \text{ contains } \overline{x_i}, \\ S & \text{otherwise.} \end{cases}$$

Then, construct f , g and h as follows:

$$\begin{aligned} f(n, s, e) &= e, \\ g(n, s, d) &= (\{S\}, \{0, 1\}, S, \\ &\quad \{S \rightarrow 0 \mid 1 \mid w_1^1 \dots w_n^1 \mid \dots \mid w_1^m \dots w_n^m \mid \underbrace{S \dots S}_{n+1} \mid \dots \mid \underbrace{S \dots S}_{2n}\}), \\ h(n, s, e') &= \begin{cases} e' & \text{if } |e'| = n, \\ \perp & \text{if } 1 < |e'| < n, \\ \top & \text{if } |e'| = 1 \text{ or } |e'| > n. \end{cases} \end{aligned}$$

For each $e \in \{0, 1\}^n$, it holds that e satisfies d iff $S \Rightarrow_{g(n, s, d)}^* f(n, s, e)$. Furthermore, for each $e' \in \{0, 1\}^*$, if $h(n, s, e') = \perp$, then $S \not\Rightarrow_{g(n, s, d)}^* e'$, because $g(n, s, d)$ generates no strings of length more than 1 and less than n ; If $h(n, s, e') = e'$, then it holds that $S \Rightarrow_{g(n, s, d)}^* e'$ iff $h(n, s, e')$ satisfies d .

Finally, consider the case that $h(n, s, e') = \top$. It is sufficient to show that, for each $k \geq 1$, it holds that $S \Rightarrow_{g(n, s, d)}^* \underbrace{S \dots S}_{kn+m}$ for each m ($1 \leq m \leq n$). If $k = 1$, then, by the definition, it holds that $S \Rightarrow_{g(n, s, d)}^* \underbrace{S \dots S}_{n+m}$ for each m ($1 \leq m \leq n$). Suppose that, for some $k \geq 1$, $S \Rightarrow_{g(n, s, d)}^* \underbrace{S \dots S}_{kn+m}$ for each m ($1 \leq m \leq n$). Then, it holds that $S \Rightarrow_{g(n, s, d)}^* \underbrace{S \dots S}_{kn+(m-1)} S \Rightarrow_{g(n, s, d)}^* \underbrace{S \dots S}_{kn+(m-1)} \underbrace{S \dots S}_{n+1} = \underbrace{S \dots S}_{(k+1)n+m}$ for each m ($1 \leq m \leq n$). Hence, $g(n, s, d)$ generates all strings of length more than n , so if $h(n, s, e') = \top$, then $S \Rightarrow_{g(n, s, d)}^* e'$. $\square$

Theorem 5. $\mathcal{L}_{\cup\text{DFA}} \trianglelefteq_{\text{pwm}} \mathcal{L}_{\text{paren}}$.

Proof. Let $M_1, \dots, M_r$ be DFAs with the same alphabet Σ ($[,] \notin \Sigma$) and with mutually distinct states. For each $M_i = (Q_i, \Sigma, \delta_i, q_0^i, F_i)$ ($1 \leq i \leq r$), construct a parenthesis grammar $G_i(n, s, M_i) = (Q_i, \Sigma \cup \{[,]\}, P_i, q_0^i)$ such that $q \rightarrow [a\delta_i(q, a)] \in P_i$ for each $q \in Q_i$ and $a \in \Sigma$; $q \rightarrow [\varepsilon]$ for each $q \in F_i$, where ε is an empty string. By using $G_i(n, s, M_i)$, let $P_{M_1, \dots, M_r}$ be the following set of productions for $S \notin (\cup_{1 \leq i \leq r} Q_i) \cup \Sigma \cup \{[,]\}$:

$$P_{M_1, \dots, M_r} = \{S \rightarrow [q_0^1] \mid \dots \mid [q_0^r]\} \cup (\cup_{1 \leq i \leq r} P_i).$$

Then, construct f , g and h as follows:

$$\begin{aligned}
f(n, s, e_1 e_2 \cdots e_l) &= [[e_1[e_2[\cdots[e_l[\varepsilon]]\cdots]]]] \text{ for } e_i \in \Sigma, \\
g(n, s, \{M_1, \dots, M_r\}) &= (\{S\} \cup (\cup_{1 \leq i \leq r} Q_i), \Sigma \cup \{[,]\}, S, P_{M_1, \dots, M_r}), \\
h(n, s, e') &= \begin{cases} e_1 \cdots e_l & \text{if } e' = [[e_1[e_2[\cdots[e_l[\varepsilon]]\cdots]]]] \text{ and } e_i \in \Sigma, \\ \perp & \text{otherwise.} \end{cases}
\end{aligned}$$

Note that $L(g(n, s, \{M_1, \dots, M_r\})) \subseteq \{[[e_1[e_2[\cdots[e_m[\varepsilon]]\cdots]]]] \mid m \geq 1, e_i \in \Sigma\}$, so if $h(n, s, e') = \perp$, then $S \not\Rightarrow_{g(n, s, \{M_1, \dots, M_r\})}^* e'$. Also it is obvious that $S \Rightarrow_{g(n, s, \{M_1, \dots, M_r\})}^* [[e_1[e_2[\cdots[e_l[\varepsilon]]\cdots]]]]$ iff $e_1 e_2 \cdots e_l \in L(M_i)$ for some i ($1 \leq i \leq r$). Hence, it holds that $\mathcal{L}_{\text{UDFA}} \leq_{\text{pwm}} \mathcal{L}_{\text{paren}}$. $\square$

Sakakibara [19] has shown that $\mathcal{L}_{\text{paren}}$ is polynomial-time learnable with membership and equivalence queries if the structural information is available. Furthermore, Sakamoto [21] has shown that $\mathcal{L}_{\text{paren}}$ is polynomial-time learnable with membership queries and *characteristic examples*. The above theorem claims that the structural information or the characteristic examples are essential for efficient learning of $\mathcal{L}_{\text{paren}}$.

4.2 Finite unions of regular pattern languages

Since each regular pattern language is regular [25], we can construct a DFA M_π such that $L(M_\pi) = L(\pi)$ for each regular pattern π as follows: Suppose that π is a regular pattern of the form $\pi = x_0 \alpha_1 x_1 \alpha_2 \cdots x_{n-1} \alpha_n x_n$, where $x_i \in X$ and $\alpha_i = a_1^i a_2^i \cdots a_{m_i}^i \in \Sigma^+$. Then, the *corresponding DFA for π* is the DFA $M_\pi = (\Sigma, Q, \delta, q_0, F)$ such that:

1. $Q = \{q_0, p_1^1, \dots, p_{m_1}^1, q_1, p_1^2, \dots, p_{m_2}^2, q_2, \dots, q_{n-1}, p_1^n, \dots, p_{m_n}^n, q_n\}$ and $F = \{q_n\}$,
2. $\delta(q_i, a) = p_1^{i+1}$ and $\delta(q_n, a) = q_n$ for each $a \in \Sigma$ and $0 \leq i \leq n-1$,
3. $\delta(p_j^i, a_j^i) = p_{j+1}^i$ and $\delta(p_{m_i}^i, a_{m_i}^i) = q_i$ for each $1 \leq i \leq n$ and $1 \leq j \leq m_i-1$,
4. $\delta(p_j^i, a) = p_1^i$ for each $a \in \Sigma$ such that $a \neq a_j^i$.

It is obvious that $|M_\pi|$ is bounded by a polynomial in $|\pi|$.

By using the corresponding DFAs, we can easily shown that $\mathcal{L}_{\text{RP}} \leq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$ by constructing the following f , g and h for each regular pattern π :

$$\begin{aligned}
f(n, s, e) &= e, \\
g(n, s, \pi) &= M_\pi, \\
h(n, s, e') &= e'.
\end{aligned}$$

Then, $\mathcal{L}_{\text{RP}}$ is polynomial-time predictable with membership queries, which is implied by the result of Matsumoto and Shinohara [15] that $\mathcal{L}_{\text{RP}}$ is polynomial-time learnable with equivalence and membership queries. Furthermore, the following theorem holds:

Theorem 6. *For each $m \geq 0$, $\mathcal{L}_{\cup_m \text{RP}} \leq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$.*

Proof. Let $\pi_1, \dots, \pi_m$ be m regular patterns. Also let $M_{\pi_i} = (Q_i, \Sigma, \delta_i, q_0^i, F_i)$ be the corresponding DFA of π_i . First, construct a DFA $M_{\pi_1, \dots, \pi_m} = (Q_1 \times \cdots \times Q_m, \Sigma, \delta, (q_0^1, \dots, q_0^m), F_1 \times \cdots \times F_m)$ such that $\delta((q_1, \dots, q_m), a) = (p_1, \dots, p_m)$ iff $\delta_i(q_i, a) = p_i$ for each i ($1 \leq i \leq m$). Then, construct f , g and h as follows:

$$\begin{aligned} f(n, s, e) &= e, \\ g(n, s, \{\pi_1, \dots, \pi_m\}) &= M_{\pi_1, \dots, \pi_m}, \\ h(n, s, e') &= e'. \end{aligned}$$

Recall that the size of $g(n, s, \{\pi_1, \dots, \pi_m\})$ is bounded by a polynomial in s , that is, $O(s^m)$. It is obvious that $L(\pi_1) \cup \dots \cup L(\pi_m) = L(M_{\pi_1, \dots, \pi_m})$, which implies that $\mathcal{L}_{\cup_m \text{RP}} \leq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$. $\square$

For the unbounded finite unions of regular pattern languages, the following theorem holds from our previous result.

Theorem 7 (Sakamoto *et al.* [22]). $\mathcal{L}_{\text{DNF}} \leq_{\text{pwm}} \mathcal{L}_{\cup \text{RP}}$.

The idea of the proof of Theorem 7 in [22] is similar as one of Theorem 4: First, for each term t_j ($1 \leq j \leq m$) in a DNF formula $d = t_1 \vee \dots \vee t_m$ over $x_1, \dots, x_n$, construct a regular pattern $\pi_j = \pi_1^j \dots \pi_n^j$ as follows:

$$\pi_i^j = \begin{cases} 1 & \text{if } t_j \text{ contains } x_i, \\ 0 & \text{if } t_j \text{ contains } \overline{x_i}, \\ x_i^j & \text{otherwise.} \end{cases}$$

Furthermore, let π be a regular pattern $x_1 \dots x_n x_{n+1}$. Then, we can construct f and h as same as Theorem 4 and g as $g(n, s, d) = \{\pi_1, \dots, \pi_m, \pi\}$. Note that $L(\pi) = \Sigma^* - \Sigma^{[n]}$, because of disallowing ε -substitutions.

For the unbounded finite unions of substring pattern languages, we obtain the following theorem.

Theorem 8. $\mathcal{L}_{\cup \text{subP}} \leq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$.

Proof. Let $\pi_1, \dots, \pi_r$ be substring patterns such that $\pi_i = x_i w_i y_i$. For a set $\{w_1, \dots, w_r\}$ of constant strings, consider the following modification of a *pattern matching machine (pmm)* $M_{w_1, \dots, w_r}$ [1]. The goto function is defined as same as a pmm. For the state j transitioned from the right-most constant of w_i by the goto function, the failure function maps j to j itself; otherwise, the failure function is defined as same as a pmm. The output function is not necessary.

Since this modified pmm is also a DFA, construct f , g and h as follows:

$$\begin{aligned} f(n, s, e) &= e, \\ g(n, s, \{\pi_1, \dots, \pi_r\}) &= M_{w_1, \dots, w_r}, \\ h(n, s, e') &= e'. \end{aligned}$$

Note that the size of $g(n, s, \{\pi_1, \dots, \pi_r\})$ is $O(|w_1| + \dots + |w_r|)$ [1]. Furthermore, it is obvious that $L(g(n, s, \{\pi_1, \dots, \pi_r\})) = L(\pi_1) \cup \dots \cup L(\pi_r)$, which implies that $\mathcal{L}_{\cup \text{subP}} \leq_{\text{pwm}} \mathcal{L}_{\text{DFA}}$. $\square$

Shinohara and Arimura [26] have discussed the inferability of $\mathcal{L}_{\cup_m \text{RP}}$, $\mathcal{L}_{\cup \text{RP}}$ and $\mathcal{L}_{\cup \text{subP}}$ in the framework of inductive inference. They have shown that $\mathcal{L}_{\cup_m \text{RP}}$ and $\mathcal{L}_{\cup \text{subP}}$ are inferable from positive data in the limit, whereas $\mathcal{L}_{\cup \text{RP}}$ is not. In contrast, by Theorem 6, 7 and 8, $\mathcal{L}_{\cup_m \text{RP}}$ and $\mathcal{L}_{\cup \text{subP}}$ are polynomial-time predictable with membership queries, whereas $\mathcal{L}_{\cup \text{RP}}$ is not if neither are DNF formulas.

5 Conclusion

In this paper, we have presented the results of pwm-reducibility on formal languages described as Fig. 1 in Section 1.

The results in Section 4.1 tell us that the efficient predictability of CFGs may be necessary to assume some deterministic application for productions. Ishizaka [10] showed that the *simple deterministic grammars* are polynomial-time learnable with *extended* equivalence and membership queries. The extended equivalence query can check whether or not the hypothesis is generated by simple deterministic grammars. It remains open whether or not simple deterministic grammars are polynomial-time predictable with membership queries.

Angluin [4] showed that $\mathcal{L}_{k\text{-bounded-CFG}}$ is polynomial-time predictable with *nonterminal* membership queries, and Sakakibara [20] extended Angluin's result to *extended simple formal systems*. In our previous work, we extended the pwm-reduction to prediction-preserving reduction with nonterminal membership queries partially [22]. It is important to investigate the properties of it in more detail.

In Section 4.2, we only deal with the finite unions of *regular* pattern languages. Many researchers such as [2, 11, 13–15, 17, 26] have developed the learnability/predictability of non-regular pattern languages. In particular, the learnability of the languages of *k-variable patterns* [2, 11] that contain at most k variables, *k μ -patterns* [15] each of which variable occurs at most k times, and *erasing patterns* [17, 26] that allow ε -substitutions have been widely studied in the various learning frameworks. It is a future work to investigate the pwm-reducibility of them or their finite unions.

Acknowledgment

The authors would like to thank Hiroki Arimura in Kyushu University for the idea of the proof of Theorem 8 and also valuable comments.

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