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by

SANJAY JAIN, FRANK STEPHAN, AND THOMAS ZEUGMANN

Division of Computer Science

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Hokkaido University
Graduate School of
Information Science and Technology

Email: thomas@ist.hokudai.ac.jp

Phone: +81-011-706-7684

Fax: +81-011-706-7684

On the Amount of Nonconstructivity in Learning Formal Languages from Text

SANJAY JAIN

*Department of Computer Science
National University of Singapore
Singapore 117417
Republic of Singapore
sanjay@comp.nus.edu.sg*

FRANK STEPHAN

*Department of Computer Science and
Department of Mathematics
National University of Singapore
Singapore 117417
fstephan@comp.nus.edu.sg*

THOMAS ZEUGMANN

*Division of Computer Science
Hokkaido University
N-14, W-9, Sapporo 060-0814, Japan
thomas@ist.hokudai.ac.jp*

Abstract

Nonconstructive computations by various types of machines and automata have been considered by e.g., Karp and Lipton [17] and Freivalds [9, 10]. They allow to regard more complicated algorithms from the viewpoint of much more primitive computational devices. The amount of nonconstructivity is a quantitative characterization of the distance between types of computational devices with respect to solving a specific problem.

This paper studies the amount of nonconstructivity needed to learn classes of formal languages. Different learning types are compared with respect to the amount of nonconstructivity needed to learn indexable classes and recursively enumerable classes, respectively, of formal languages from positive data. Matching upper and lower bounds for the amount of nonconstructivity needed are shown.

1. Introduction

The research subject studied in this paper derives its motivation from various sources which we shortly present below. Nonconstructive methods of proof in mathematics have a rather long and dramatic history. The debate was especially passionate when mathematicians tried to overcome the crisis concerning the foundations of mathematics.

The situation changed slightly in the forties of the last century, when nonconstructive methods found their way to discrete mathematics. In particular, Paul Erdős used nonconstructive proofs masterly, beginning with the paper [8].

Another influential paper was Bārzdīņš [3], who introduced the notion of advice in the setting of Kolmogorov complexity of recursively enumerable sets. Karp and Lipton [17] introduced the notion of a Turing machine that takes advice to understand under what circumstances nonuniform upper bounds can be used to obtain uniform upper bounds. Damm and Holzer [7] adapted the notion of advice for finite automata. Later Cook and Krajiček [6] initiated the study of proof systems that use advice for the verification of proofs. Even more recently, Beyersdorff *et al.* [4] continued along this line of research.

Quite often, we experience that finding a proof for a new deep theorem is triggered by a certain amount of inspiration. Being inspired does not mean that we do not have to work hard in order to complete the proof and to elaborate all the technical details. However, this work is quite different from enumerating all possible proofs until we have found the one sought for. Also, as experience shows, the more complicated the proof, the higher is the amount of inspiration needed. These observations motivated Freivalds [9, 10] to introduce a qualitative approach to measure the amount of nonconstructivity (or advice) in a proof. Analyzing three examples of nonconstructive proofs led him to a notion of nonconstructive computation which can be used for many types of automata and machines and which essentially coincides with Karp and Lipton's [17] notion when applied to Turing machines.

As outlined by Freivalds [9, 10], there are several results in the theory of inductive inference of recursive functions which suggest that the notion of nonconstructivity may be worth a deeper study in this setting, too. Subsequently, Freivalds and Zeugmann [11] introduced a model to study the amount of nonconstructivity needed to learn recursive functions.

In the present paper we generalize the model of Freivalds and Zeugmann [11] to the inductive inference of formal languages. That is, we aim to characterize the difficulty to learn classes of formal languages from positive data by using the *amount of nonconstructivity* needed to learn these classes. We shortly describe this model. The learner receives, as usual, growing initial segments of a text for the target language L , where a text is any infinite sequence of strings and a special pause symbol $\#$ such that the range of the text minus the pause symbol contains all elements of L and nothing else. In addition, the learner receives as a second input a bitstring of finite length which we call *help-word*. If the help-word is correct, the learner learns in the desired sense. Since there are infinitely many languages to learn, a parameterization is necessary, i.e., we allow for every n a possibly different help-word and we require the learner to learn every language contained in $\{L_0, \dots, L_n\}$ with respect to the hypothesis space $(L_i)_{i \in \mathbb{N}}$ chosen (cf. Definition 6). The difficulty of the learning problem is then measured by the length of the help-words needed, i.e., in terms of the growth rate of the function d bounding this length. As in previous approaches, the help-word does *not* just provide

an answer to the learning problem. There is still much work to be done by the learner.

First, we consider the learnability of indexable classes in the limit from positive data and ask for the amount of nonconstructivity needed to learn them. This is quite a natural choice, since even simple indexable subclasses of the class of all regular languages are known *not* to be inferable in the limit from positive data (cf. [13, 15, 22]). Second we investigate the amount of nonconstructivity needed to infer recursively enumerable classes of recursively enumerable languages. Moreover, several variations of Gold's [13] model of learning in the limit have been considered (cf., e.g., [15, 20] and the references therein). Thus, it is only natural to consider some of these variations, too. In particular, we shall study *conservative* learning and *strong-monotonic* inference.

We then prove upper and lower bounds for the amount of nonconstructivity in learning classes of formal languages from positive data. The usefulness of this approach is nicely reflected by our results which show that the function d may considerably vary. In particular, the function d may be arbitrarily slow growing for learning indexable classes in the limit from positive data (cf. Theorem 1), while we have an upper bound of $\log n$ and a lower bound of $\log n - 2$ for conservative learning of indexable classes from positive data (cf. Theorems 2 and 3). Furthermore, we have a $2 \log n$ upper bound and a $2 \log n - 4$ lower bound for strong-monotonic inference of indexable classes from positive data (cf. Theorems 4 and 5).

Moreover, the situation changes considerably when looking at recursively enumerable classes of recursively enumerable languages. For learning in the limit from positive data we have an upper bound of $\log n$ and a lower bound of $\log n - 2$, while for conservative learning even any limiting recursive bound on the growth of the function d is not sufficient to learn all recursively enumerable classes of recursively enumerable languages from positive data (cf. Theorems 7, 8 and 9).

2. Preliminaries

Any unspecified notations follow Rogers [21]. In addition to or in contrast with Rogers [21] we use the following. By $\mathbb{N} = \{0, 1, 2, \dots\}$ we denote the set of all natural numbers, and we set $\mathbb{N}^+ = \mathbb{N} \setminus \{0\}$.

The cardinality of a set S is denoted by $|S|$. We write $\wp(S)$ for the power set of set S . Let \emptyset , \in , \subset , \subseteq , \supset , and \supseteq denote the empty set, element of, proper subset, subset, proper superset, and superset, respectively. Let S_1, S_2 be any sets; then we write $S_1 \Delta S_2$ to denote the symmetric difference of S_1 and S_2 , i.e., $S_1 \Delta S_2 = (S_1 \setminus S_2) \cup (S_2 \setminus S_1)$. By $\max S$ and $\min S$ we denote the maximum and minimum of a set S , respectively, where, by convention, $\max \emptyset = 0$ and $\min \emptyset = \infty$.

We use \mathfrak{T} to denote the set of all total functions of one variable over \mathbb{N} . Let $n \in \mathbb{N}^+$; then the set of all partial recursive functions and of all recursive functions of one and n variables over \mathbb{N} is denoted by \mathcal{P} , \mathcal{R} , \mathcal{P}^n , \mathcal{R}^n , respectively. Furthermore,

let $f \in \mathcal{P}$, then we use $\text{dom}(f)$ to denote the *domain* of the function f , that is, $\text{dom}(f) = \{x \mid x \in \mathbb{N}, f(x) \text{ is defined}\}$. By $\text{range}(f)$ we denote the *range* of f , i.e., $\text{range}(f) = \{f(x) \mid x \in \text{dom}(f)\}$.

It is technically most convenient to define recursively enumerable families of recursively enumerable languages as follows. Any $\psi \in \mathcal{P}^2$ is called a numbering. Let $\psi \in \mathcal{P}^2$, then we write ψ_i instead of $\lambda x.\psi(i, x)$. Furthermore, we set $W_i^\psi = \text{dom}(\psi_i)$ and refer to it as the i th enumerated language. Clearly, the sets $W_i^\psi \subseteq \mathbb{N}$ are recursively enumerable.

A numbering $\varphi \in \mathcal{P}^2$ is called a *Gödel numbering* for \mathcal{P} (cf. Rogers [21]) if $\mathcal{P}_\varphi = \mathcal{P}$, and for every numbering $\psi \in \mathcal{P}^2$, there is a *compiler* $c \in \mathcal{R}$ such that $\psi_i = \varphi_{c(i)}$ for all $i \in \mathbb{N}$. If $\varphi \in \mathcal{P}^2$ is any arbitrarily fixed Gödel numbering then we also use W_i as a shorthand for W_i^φ .

A function $f \in \mathcal{P}$ is said to be *strictly monotonic* provided for all $x, y \in \mathbb{N}$ with $x < y$ we have, if both $f(x)$ and $f(y)$ are defined then $f(x) < f(y)$. By \mathcal{R}_{mon} we denote the set of all strictly monotonic recursive functions.

Let Σ be any fixed finite alphabet, and let Σ^* be the free monoid over Σ . Any $L \subseteq \Sigma^*$ is a language. Furthermore, we fix a symbol $\#$ such that $\# \notin \Sigma$. We denote the empty string by λ and use $|w|$ to denote the length of any string $w \in \Sigma^*$. By \mathcal{REG} we denote the class of all *regular* languages (cf., e.g., [23]). Furthermore, we use \mathcal{C} or \mathcal{L} to denote any (infinite) class and family of languages, respectively.

Definition 1 (Gold [13]). *Let L be any language. Every total function $t: \mathbb{N} \rightarrow \Sigma^* \cup \{\#\}$ with $\{t(j) \mid j \in \mathbb{N}\} \setminus \{\#\} = L$ is called a text for L .*

Note that the symbol $\#$ denotes pauses in the presentation of data. Furthermore, there is no requirement concerning the computability of a text. So, any order and any number of repetitions is allowed. For any $n \in \mathbb{N}$ we use $t[n]$ to denote the *initial segment* $(t(0), \dots, t(n))$. Additionally we use $\text{content}(t[n]) =_{\text{df}} \{t(0), \dots, t(n)\} \setminus \{\#\}$ and $\text{content}(t) =_{\text{df}} \{t(j) \mid j \in \mathbb{N}\} \setminus \{\#\}$ to denote the content of an initial segment and of a text, respectively.

An algorithmic *learner* M finds a rule (grammar) from growing initial segments of a text. On each initial segment the learner M has to output a hypothesis which is a natural number, i.e., $M(t[n]) \in \mathbb{N}$. Then the sequence $(M(t[n]))_{n \in \mathbb{N}}$ has to *converge* (to some representation of the input), i.e., there is a $j \in \mathbb{N}$ such that $M(t[n]) = j$ for all but finitely many $n \in \mathbb{N}$.

So, we still have to specify the semantics of the numbers output by M . In order to do so, we need the following.

Definition 2 (Angluin [2]). *A family $(L_j)_{j \in \mathbb{N}}$ of languages is said to be uniformly recursive if there exists a recursive function $f: \mathbb{N} \times \Sigma^* \rightarrow \{0, 1\}$ such that $L_j = \{w \mid w \in \Sigma^*, f(j, w) = 1\}$ for all $j \in \mathbb{N}$. We refer to f as a decision function.*

Definition 3. A class \mathcal{C} of non-empty recursive languages is said to be indexable if there is a family $(L_j)_{j \in \mathbb{N}}$ of uniformly recursive languages such that $\mathcal{C} = \{L_j \mid j \in \mathbb{N}\}$. Such a family is said to be an indexing of \mathcal{C} .

By \mathcal{ID} we denote the collection of all indexable classes.

Note that \mathcal{REG} is an indexable class. Also, the class of all context-free languages and the class of all context-sensitive languages form an indexable class. Further information concerning indexable classes and their learnability can be found in [20, 22].

So, when dealing with the learnability of indexable classes, it is only natural to interpret the hypotheses output by M with respect to a chosen indexing of a class containing the target class \mathcal{C} (cf. Definition 4 below). On the other hand, when considering recursively enumerable classes \mathcal{C} of recursively enumerable languages then we always take as hypothesis space the family $(W_i^\psi)_{i \in \mathbb{N}}$, where $\psi \in \mathcal{P}^2$ is the numbering defining the class \mathcal{C} .

Definition 4. Let \mathcal{C} be an indexable class. A family $\mathcal{H} = (L_j)_{j \in \mathbb{N}}$ is said to be an indexed hypothesis space for \mathcal{C} if $(L_j)_{j \in \mathbb{N}}$ is uniformly recursive and $\mathcal{C} \subseteq \{L_j \mid j \in \mathbb{N}\}$.

Following Lange and Zeugmann [19], if $\mathcal{C} = \{L_j \mid j \in \mathbb{N}\}$ then we call \mathcal{H} *class preserving* and if $\mathcal{C} \subseteq \{L_j \mid j \in \mathbb{N}\}$ then the hypothesis space \mathcal{H} is said to be *class comprising*.

Now we are ready to provide the formal definition of learning in the limit from text. Following Gold [13] we call our learners *inductive inference machines* (abbr. IIM). To unify notations, in the definitions below we use $\mathcal{H} = (h_j)_{j \in \mathbb{N}}$ to denote our hypothesis spaces, where we assume the interpretation given above.

Definition 5 (Gold [13]). Let \mathcal{C} be any class of languages, let $\mathcal{H} = (h_j)_{j \in \mathbb{N}}$ be a hypothesis space for \mathcal{C} , and let $L \in \mathcal{C}$. An IIM M is said to learn L in the limit from text with respect to \mathcal{H} if

- (1) for every text t for L there is a $j \in \mathbb{N}$ such that the sequence $(M(t[n]))_{n \in \mathbb{N}}$ converges to j , and
- (2) $L = h_j$.

An IIM M learns \mathcal{C} in the limit from text with respect to \mathcal{H} if M learns all $L \in \mathcal{C}$ in the limit from text with respect to \mathcal{H} .

The collection of all classes \mathcal{C} for which there is an IIM M and a hypothesis space \mathcal{H} such that M learns \mathcal{C} in the limit from text with respect to \mathcal{H} is denoted by LimTxt .

Since, by the definition of convergence, only finitely many data of L were seen by the IIM upto the (unknown) point of convergence, whenever an IIM learns the possibly infinite language L , indeed some form of learning must have taken place. For this reason, hereinafter the terms *learn*, *infer*, and *identify* are used interchangeably.

In the following modifications of Definition 5 additional requirements are made. An IIM M is said to be *consistent* if for all relevant texts t and all $n \in \mathbb{N}$ the condition $\text{content}(t[n]) \subseteq h_{M(t[n])}$ is satisfied (cf. Angluin [1]).

Following Angluin [2], an IIM M is said to be *conservative* if for all relevant texts t and all $n, m \in \mathbb{N}$ the following condition is satisfied. If $j = M(t[n]) \neq M(t[n+m])$ then $\text{content}(t[n+m]) \not\subseteq h_j$.

We call an IIM M *strong-monotonic* if for all relevant texts t and all $n, m \in \mathbb{N}$ the following condition is satisfied. If $j = M(t[n]) \neq M(t[n+m]) = k$ then $h_j \subseteq h_k$ (cf. Jantke [16], Lange and Zeugmann [18]).

We denote the resulting learning types by *ConsTxt*, *ConsvTxt*, and *SmonTxt*, respectively.

After having defined several learning models, it is only natural to ask why should we study learning with nonconstructivity. The answer is given by the fact that many interesting language classes are *not learnable from text*. As shown in [22], even quite simple classes cannot be learned from text, e.g., the class

$$\mathcal{C} = (L_j)_{j \in \mathbb{N}}, \text{ where } L_0 = \{a^j \mid j \in \mathbb{N}^+\} \text{ and } L_j = \{a^\ell \mid 1 \leq \ell \leq j\}. \quad (1)$$

We aim to characterize quantitatively the difficulty of such learning problems by measuring the amount of nonconstructivity needed to solve them.

The learners used for *nonconstructive* inductive inference take as input not only growing initial segments $t[n]$ of a text t but also a *help-word* w . The help-words are assumed to be encoded in binary. So, for such learners we write $M(t[n], w)$ to denote the hypothesis output by M . Then, for all the learning types defined above, we say that M nonconstructively identifies L with the help-word w provided that for every text t for L the sequence $(M(t[n], w))_{n \in \mathbb{N}}$ converges to a number j such that $h_j = L$ (for *LimTxt*) and M is consistent (conservative, strong-monotonic) for *ConsTxt* (for *ConsvTxt*, and *SmonTxt*), respectively. More formally we have the following definition.

Definition 6. *Let \mathcal{C} be any class of languages, let $\mathcal{H} = (h_j)_{j \in \mathbb{N}}$ be a hypothesis space for \mathcal{C} , and let $d \in \mathcal{R}$. An IIM M infers \mathcal{C} with nonconstructivity $d(n)$ in the limit with respect to \mathcal{H} , if for each $n \in \mathbb{N}$ there is a help-word w of length at most $d(n)$ such that for every $L \in \mathcal{C} \cap \{h_0, h_1, \dots, h_n\}$ and every text t for L the sequence $(M(t[m], w))_{m \in \mathbb{N}}$ converges to a hypothesis j satisfying $h_j = L$.*

Clearly, Definition 6 can be directly modified to obtain *nonconstructive conservative and strong-monotonic learning*.

Looking at Definition 6 it should be noted that the IIM may need to know either an appropriate upper bound for n or even the precise value of n in order to exploit the fact that the target language L is from $\mathcal{C} \cap \{h_0, h_1, \dots, h_n\}$.

To simplify notation, we make the following convention. Whenever we talk about nonconstructivity $\log n$, we assume that the logarithmic function to the base 2 is replaced by its integer valued counterpart $\lfloor \log n \rfloor + 1$, where $\log 0 =_{df} 1$.

Now we are ready to present our results. Note that some proofs have been influenced by ideas developed in the quite a different context, i.e., the paradigm of learning by erasing (also called co-learning). We do not explain it here but refer the reader to Jain *et al.* [14] as well as to Freivalds and Zeugmann [12].

3. Results

Already Gold [13] showed that $\mathcal{REG} \notin \text{LimTxt}$ and as mentioned in (1), even quite simple subclasses of \mathcal{REG} are not in LimTxt . So, we start our investigations by asking for the *amount of nonconstructivity* needed to identify any indexable class in the limit from text with respect to any indexed hypothesis space \mathcal{H} .

3.1. Nonconstructive Learning of Indexable Classes

As we shall see, the needed amount of nonconstructivity is surprisingly small. In order to show this result, for every function $d \in \mathcal{R}_{mon}$ we define its *inverse* d_{inv} as follows $d_{inv}(n) = \mu y[d(y) \geq n]$ for all $n \in \mathbb{N}$. Recall that $\text{range}(d)$ is recursive for all $d \in \mathcal{R}_{mon}$. Thus, for all $d \in \mathcal{R}_{mon}$ we can conclude that $d_{inv}(n) \in \mathcal{R}$.

Theorem 1. *Let $\mathcal{C} \in \mathcal{ID}$ be arbitrarily fixed, let $d \in \mathcal{R}_{mon}$ be any function, and let $\mathcal{H} = (L_j)_{j \in \mathbb{N}}$ be any indexed hypothesis space for \mathcal{C} . Then there is a computable IIM M such that the class \mathcal{C} can be identified with nonconstructivity $\log d_{inv}(n)$ in the limit from text with respect to \mathcal{H} .*

Proof. Assuming any help-word w of length precisely $\log d_{inv}(n)$, the IIM M creates a bitstring containing only 1s that has the same length as w . This bitstring is interpreted as a natural number k .

So, $k \geq d_{inv}(n)$, and thus

$$u_* =_{df} d(k) \geq d(d_{inv}(n)) \geq n . \quad (2)$$

We continue to define the IIM M in a way such that it will learn every language $L \in \mathcal{C} \cap \{L_0, \dots, L_{u_*}\}$ from every of its texts. So, fix any such L , let t be any text for L , and let $m \in \mathbb{N}$.

Now, the idea to complete the proof is as follows. In the limit, the IIM M can determine the *number* ℓ of different languages enumerated in L_0, \dots, L_{u_*} as well as the *least indices* j_1, \dots, j_ℓ of them and can then find the language among them which is equal to L . We assume the lexicographical ordering \leq_{lo} of all strings from Σ^* , i.e., $s_i \leq_{lo} s_{i+1}$ for all $i \in \mathbb{N}$.

Using m , $t[m]$, and the decision function f for \mathcal{H} , the IIM M computes the least number r such that $m \leq r$ and $s \leq_{lo} s_r$ for all $s \in \text{content}(t[m])$. Next, M computes

$$\begin{aligned} L_0^r &= \{w \mid w \leq_{lo} s_r, f(0, w) = 1\} \\ L_1^r &= \{w \mid w \leq_{lo} s_r, f(1, w) = 1\} \\ &\cdot \\ &\cdot \\ &\cdot \\ L_{u_*}^r &= \{w \mid w \leq_{lo} s_r, f(u_*, w) = 1\}, \end{aligned}$$

and chooses the least indices j_1, \dots, j_{ℓ_m} from $0, 1, \dots, u_*$ of *all* the distinct languages in the list $L_0^r, \dots, L_{u_*}^r$. From these languages $L_{j_z}^r$ all those are deleted for which $\text{content}(t[m]) \not\subseteq L_{j_z}^r$ (the inconsistent ones). From the remaining indices, the least index j is output such that $|L_j^r \setminus \text{content}(t[m])|$ is minimal.

Now, it is easy to see that the sequence $(\ell_m)_{m \in \mathbb{N}}$ converges to ℓ , the number of the different languages enumerated in L_0, \dots, L_{u_*} , and that the IIM M finds in the limit the least indices j_1, \dots, j_ℓ for these pairwise different languages.

From these languages $L_{j_1}, \dots, L_{j_\ell}$ the ones satisfying $L \setminus L_{j_z} \neq \emptyset$ are deleted.

That leaves all those L_{j_z} with $L \subseteq L_{j_z}$. Now, by assumption there is a least $j \in \{0, \dots, u_*\}$ with $L_j = L$. If $L \subset L_{j_z}$, then there is a string $s \in L_{j_z} \setminus L$, and as soon as this string appears in the competition, the index j wins. Thus, the sequence $(M(t[m], w))_{m \in \mathbb{N}}$ converges to j . \square

So there is *no smallest* amount of nonconstructivity needed to learn \mathcal{REG} and any subset thereof in the limit from text. But the amount of nonconstructivity cannot be zero, since then we would have $\mathcal{REG} \in \text{LimTxt}$. One can define a total function $t \in \mathfrak{T}$ such that $t(n) \geq d(n)$ for all $d \in \mathcal{R}_{\text{mon}}$ and all but finitely many n . Consequently, $\log t_{\text{inv}}$ is then a lower bound for the amount of nonconstructivity needed to learn \mathcal{REG} in the limit from text for the technique used to show Theorem 1.

We continue by asking what amount of nonconstructivity is needed to obtain *conservative* learning from text for any indexable class. Now, the situation is intuitively more complex, since $\text{ConsvTxt} \subset \text{LimTxt}$ (cf. [2, 19]). Also, it is easy to see that the IIM M given in the proof of Theorem 1 is in general *not* conservative. But the basic idea still works *mutatis mutandis* provided we know the number ℓ of different languages enumerated in L_0, \dots, L_n .

Theorem 2. *Let $\mathcal{C} \in \mathcal{ID}$ be arbitrarily fixed, and let $\mathcal{H} = (L_j)_{j \in \mathbb{N}}$ be an indexed hypothesis space for \mathcal{C} . Then there is a computable IIM M such that the class \mathcal{C} can be conservatively identified with nonconstructivity $\log n$ from text with respect to \mathcal{H} .*

Proof. Let $\mathcal{H} = (L_j)_{j \in \mathbb{N}}$ be any indexed hypothesis space for \mathcal{C} , and let $n \in \mathbb{N}$. The help-word w is defined as follows. Since the IIM also needs to know a bound on n , we always assume n to be a power of 2. Intuitively, we then add one bit and write the

binary representation of the *exact number* ℓ of pairwise different languages enumerated in L_0, \dots, L_n behind the leading 1 including leading zeros. But of course we do not need the leading 1 in the help-word, since it can be added by the IIM M . So if the help-word w has length k , then the added leading 1 with $k - 1$ zeros gives n and the bitstring w without the added leading 1 gives ℓ .

Given ℓ , the desired IIM M can find the least indices of these ℓ pairwise different languages by using the decision function f from the proof of Theorem 1 above, where we additionally require r to be large enough to detect ℓ different languages.

The rest is done inductively. The IIM M checks whether or not $t(0) \in L_{j_z}^r$, where $z = 1, \dots, \ell$, and deletes all languages which fail. Then M orders the remaining sets $L_{j_z}^r$ with respect to set inclusion, and outputs the index of the minimal one with the smallest index. For $m > 0$, M checks whether or not $\text{content}(t[m]) \subseteq L_{M(t[m-1])}$. If it is, it outputs $M(t[m-1])$.

Otherwise, it checks whether or not $\text{content}(t[m]) \subseteq L_{j_z}^r$, $z = 1, \dots, \ell$, and deletes all languages which fail. Then M orders the remaining sets $L_{j_z}^r$ with respect to set inclusion, and outputs the index of the minimal one with the smallest index. \square

We also have the following lower bound.

Theorem 3. *There is a class $\mathcal{C} \in \mathcal{ID}$ and an indexed hypothesis space \mathcal{H} for it such that for every IIM that learns \mathcal{C} conservatively with respect to \mathcal{H} less than $\log n - 2$ many bits of nonconstructivity are not enough.*

Proof. This proof is done by diagonalization. We have to define a family $(L_j)_{j \in \mathbb{N}}$ of uniformly recursive languages and then to show that it satisfies the assertion of the theorem. Let $\varphi \in \mathcal{P}^2$ be any fixed Gödel numbering of \mathcal{P} . We interpret every partial recursive function φ_i as an IIM, i.e., we set $M_i = \varphi_i$. Below we shall use the alphabet $\Sigma = \{a, b\}$ and assume any fixed recursive encoding $\langle \cdot, \cdot \rangle$ of all finite sequences from $\Sigma^* \times \{0, 1\}^*$ onto the natural numbers. To simplify notation we shortly write $M(\sigma, \beta)$ instead of $M(\langle \sigma, \beta \rangle)$, where σ is any finite sequence of elements from Σ^* and $\beta \in \{0, 1\}^*$. Also, we write $M(\sigma, \beta) \downarrow = \ell$ provided the computation of $M(\sigma, \beta)$ stops in a finite number of steps and outputs ℓ .

Our construction is organized in blocks, where in Block i , $i \in \mathbb{N}$, we shall define languages L_{2^i+j} for all $0 \leq j \leq 2^i - 1$, i.e., the languages $L_{2^i}, \dots, L_{2^{i+1}-1}$ by diagonalizing against M_{i-1} and all possible help-words $\beta \in \{0, 1\}^*$ with $0 \leq |\beta| < i$ including the empty one λ . Note that there are $2^i - 1$ many such β .

Part of the construction is based on an idea developed by Angluin [2]. For the sake of completeness we set $L_0 = \{b\}$ and define $L_1 = \{bb\}$.

For $i \geq 1$ we have to define 2^i many languages. In addition, we also define a text t_i incrementally. This is done in Block i as follows.

For $j = 0, \dots, 2^i - 1$ execute the following.

Initialize $L_{2^i+j} := \{a^i\}$, set $C_i = \{\beta \mid \beta \in \{0, 1\}^*, 0 \leq |\beta| < i\}$, $D_i = \emptyset$, and initialize the text t_i by defining $t_i(0) = a^i$.

Execute Step 1.

Step k . Execute Instruction (A).

(A) Let $t_i(k) = a^i b^k$. Execute Instruction (B).

(B) Let $L_{2^i+j} := L_{2^i+j} \cup \{a^i b^k\}$ for all $j \in \{0, \dots, 2^i - 1\} \setminus D_i$.
Execute Instruction (C).

(C) For all $\beta \in C_i$ and all $\ell = 0, 1, \dots, k$ check whether or not

$$M_{i-1}(t_i[\ell], \beta) \downarrow = 2^i + j \quad \text{for some } j \in \{0, \dots, 2^i - 1\} \setminus D_i \quad (3)$$

can be verified within at most k steps of computation until the first such β, j and ℓ are found, if ever.

If β and j are found then let $j_* := \min(\{0, \dots, 2^i - 1\} \setminus (D_i \cup \{j\}))$;
 $C_i := C_i \setminus \{\beta\}$ and $D_i := D_i \cup \{j_*\}$.

Else do nothing.

Goto Step $k + 1$.

Now, it is easy to see that $(L_\ell)_{\ell \in \mathbb{N}}$ is a family of uniformly recursive languages.

Let $i \geq 1$ be arbitrarily fixed and suppose by way of contradiction that M_{i-1} conservatively learns all the languages L_{2^i+j} , where $0 \leq j \leq 2^i - 1$, defined in Block i by using help-words of size less than i . Note that $i - 1 = \lfloor \log(2^{i+1} - 1) \rfloor + 1 - 2 = \log(2^{i+1} - 1) - 2$ by the convention made at the very end of the Preliminaries.

Note that in Instruction (C), if β and j have been found, then exactly one element is added to D_i . Since there are only $2^i - 1$ possible help-words β , we conclude that the set $\{0, 1, 2, \dots, 2^i - 1\} \setminus D_i$ is never empty. So, there is always a number j_* and by construction $j \neq j_*$.

Let β be any of the admissible help-words. We distinguish the following cases.

Case 1. The help-word β is not removed from C_i in any of the steps.

In this case M_{i-1} fails to identify L_{2^i+j} for all $j \in \{0, 1, 2, \dots, 2^i - 1\} \setminus D_i$ when using the help-word β . Note that the D_i used here is the ‘‘eventual value of D_i .’’ Since $\{0, 1, 2, \dots, 2^i - 1\} \setminus D_i \neq \emptyset$, at least one language is not learned, a contradiction.

Case 2. The help-word β is removed from C_i in some Step k .

Let j and j_* be as described in Instruction (C) above when executing this Step k . Then by construction we know that $\text{content}(t_i[k]) = L_{2^i+j_*}$, since j_* is added to D_i and so $L_{2^i+j_*}$ is not changed any further in Instruction (B). On the other hand, j is still not in D_i and thus $L_{2^i+j_*} \subset L_{2^i+j}$ after the execution of step $k + 1$. Consequently, the IIM M_{i-1} cannot learn $L_{2^i+j_*}$ conservatively when using the help-word β , a contradiction.

Therefore, for every help-word, at least one of the languages in Block i is not learned by M_{i-1} or on which M_{i-1} is not conservative.

Finally, suppose that there is an IIM M that learns the indexable class $(L_\ell)_{\ell \in \mathbb{N}}$ with nonconstructivity $\log n - 2$ with respect to the hypothesis space $(L_\ell)_{\ell \in \mathbb{N}}$. Then there must be an $i \in \mathbb{N}$ such that $M = M_i$. Consequently, for $n = 2^{i+2} - 1$ there must be a help-word w of length less than $i + 1$ such that M_i conservatively learns all the languages $L_0, \dots, L_{2^{i+2}-1}$ from any of its texts and w .

Now, we consider in particular Block $i + 1$, i.e., the languages $L_{2^{i+1}}, \dots, L_{2^{i+2}-1}$. Since $|w| < i + 1$ there must be a β such that $w = \beta$ for some $\beta \in C_{i+1}$. But as our discussion above showed, this is impossible. So, we have a contradiction and our supposition must be false. Hence, the theorem is shown. \square

Finally, we look at strong-monotonic learning. Again the situation is more complex, since $SmonTxt \subset ConsvTxt$ (cf. [19]). We also add $L_0 = \emptyset$ to every hypothesis space allowed, i.e., we always consider class comprising hypothesis spaces.

Theorem 4. *Let $\mathcal{C} \in \mathcal{ID}$ be arbitrarily fixed, and let $\mathcal{H} = (L_j)_{j \in \mathbb{N}}$ be an indexed hypothesis space for \mathcal{C} . Then there is a computable IIM M such that the class \mathcal{C} can be strong-monotonically identified with nonconstructivity $2 \log n$ from text with respect to \mathcal{H} .*

Proof. The key observation is that it suffices to know the following number

$$p = |\{(i, j) \mid L_i \not\subseteq L_j, i, j = 0, \dots, n\}|.$$

So, the help-word is just the binary encoding of p and n which is done *mutatis mutandis* as in the proof of Theorem 2. The rest is not too difficult, and thus omitted. \square

Again, the bound given in Theorem 4 cannot be improved substantially, since we have the following lower bound.

Theorem 5. *There is a class $\mathcal{C} \in \mathcal{ID}$ and an indexed hypothesis space \mathcal{H} for it such that for every IIM that learns \mathcal{C} strong-monotonically with respect to \mathcal{H} less than $2 \log n - 4$ many bits of nonconstructivity are not enough.*

Proof. Again we use diagonalization to show the desired result. We construct a family $(L_j)_{j \in \mathbb{N}}$ of uniformly recursive languages. As in the proof of Theorem 3 we set $M_i = \varphi_i$, where $\varphi \in \mathcal{P}^2$ is any fixed Gödel numbering of \mathcal{P} , and we use the same notations as there for IIMs. The construction is done in Blocks, where in Block i , $i \in \mathbb{N}$, we shall define languages L_{2^i+j} for all $j \in \{0, \dots, 2^i - 1\}$ by diagonalizing against M_{i-1} and the length-lexicographically first $2^{2^{i-1}}$ possible help-words $\beta \in \{0, 1\}^*$. For the sake of completeness we set $L_0 = \emptyset$ and $L_1 = \{b\}$.

For every $i \geq 1$ we have to define 2^i many languages. We divide these languages in two groups of size 2^{i-1} , where the first group contains the languages L_{2^i+j} for

$j = 0, \dots, 2^{i-1} - 1$. In order to simplify notation, we refer to the languages in the first group as L_{p_j} , $j = 0, \dots, 2^{i-1} - 1$, i.e., $L_{p_j} = L_{2^i+j}$. The second group contains the remaining languages L_{2^i+j} for $j = 2^{i-1}, \dots, 2^i - 1$. We refer to these languages as L_{q_k} , where $k = 0, \dots, 2^i - 1$, i.e., $L_{q_k} = L_{2^i+2^{i-1}+k}$. Since the construction is a bit more complicated, for each L_{p_j} we incrementally define a text t_i^j , where $j = 0, \dots, 2^{i-1} - 1$.

In Block i do the following.

Initialize the languages $L_{p_j} := \{a^i\}$ for all $j = 0, \dots, 2^{i-1} - 1$ and $L_{q_k} := \{a^i\}$ for all $k = 0, \dots, 2^i - 1$, set $C_i = \{\beta \mid 0 \leq |\beta| < 2(i-1)\} \cup \{0^{2(i-1)}\}$, and let $t_i^j(0) := a^i$ for all $j = 0, \dots, 2^{i-1} - 1$. Furthermore, set $x = 0$ and set $D_j = \emptyset$.

Start Step 1.

Step s . Execute the following loop.

For $j = 0$ to $2^{i-1} - 1$ do

$x := x + 1$;

$L_{p_j} := L_{p_j} \cup \{a^i b^x\}$;

$t_i^j(x) := a^i b^x$;

$t_i^z(x) := t_i^z(x-1)$ for all $z \in \{0, \dots, 2^{i-1} - 1\} \setminus \{j\}$;

$L_{q_z} := L_{q_z} \cup \{a^i b^x\}$ for all $z = 0, \dots, 2^i - 1$;

For $k = 0$ to $2^{i-1} - 1$ do

If $k \in D_j$ **pass**;

For all $\beta \in C_i$ and $\ell = 0, \dots, x$ check whether or not

$$M_{i-1}(t_i^j[\ell], \beta) \downarrow = p_j \tag{4}$$

can be verified within at most s steps of computation.

If the test (4) is satisfied for a $\beta \in C_i$ then

$C_i := C_i \setminus \{\beta\}$;

$x := x + 1$;

$L_{p_j} := L_{p_j} \cup \{a^i b^x\}$;

$t_i^j(x) := a^i b^x$;

$t_i^z(x) := t_i^z(x-1)$ for all $z \in \{0, \dots, 2^{i-1} - 1\} \setminus \{j\}$;

$L_{q_z} := L_{q_z} \cup \{a^i b^x\}$ for all $z \in \{0, \dots, 2^i - 1\} \setminus \{k\}$;

$D_j = D_j \cup \{k\}$;

Else do nothing.

end

end

If $C_i \neq \emptyset$ start Step $k + 1$.

Note that the instruction `pass` in the second loop is used to indicate that in the case that $k \in D_j$ nothing has to be done. That is, the program continues the loop with $k + 1$ if $k + 1 \leq 2^{i-1} - 1$; otherwise it returns to the first loop.

It should also be noted that we introduced the variable x above just to make sure that all texts have the same length and that any new element added to the corresponding languages has not been inserted in an earlier execution of the nested loops in the current step nor did we insert the new element in a previous step.

Now, it is easy to see that $(L_\ell)_{\ell \in \mathbb{N}}$ is a family of uniformly recursive languages.

Let $i \geq 1$ be arbitrarily fixed and suppose by way of contradiction that M_{i-1} strong-monotonically learns all the languages L_{2^i+j} , where $0 \leq j \leq 2^i - 1$, defined in Block i by using help-words of size less than $2(i-1)$ or the help word $0^{2(i-1)}$. Note that $2(i-1) = 2(\lfloor \log(2^{i+1} - 1) \rfloor + 1) - 4 = 2 \log(2^{i+1} - 1) - 4$ by the convention made at the very end of the Preliminaries.

Furthermore, note that by construction for every j the set D_j can be updated at most 2^{i-1} many times. As soon as $D_j = \{0, \dots, 2^{i-1} - 1\}$ the second loop does nothing for this j . Also note that for any change of some D_j a success in Test (4) is necessary. Thus, when $D_j = \{0, \dots, 2^{i-1} - 1\}$ for all j , we must have also deleted all help-words.

Let β be any of the admissible help-words. We distinguish the following cases.

Case 1. The help-word β is not removed from C_i in any of the steps.

In this case, there must be at least one j_* such that $D_{j_*} \subset \{0, \dots, 2^{i-1} - 1\}$. So, if the Test (4) does never succeed for such a j_* , when using the help-word β then for $L_{p_{j_*}}$ we have a text $t_i^{j_*}$ on which the IIM M_{i-1} fails to learn $L_{p_{j_*}}$ when using β . So, the IIM M_{i-1} does not learn $L_{p_{j_*}}$ when using β as help-word for at least one text, a contradiction.

Case 2. The help-word β is removed from C_i in some step.

Let s be the step in which this happens. Then there must be a j_* and a k_* such that the variable j in the first loop takes the value j_* and the variable k in the second loop takes the value k_* and the Test (4) is satisfied for some ℓ . This also implies that $k_* \notin D_{j_*}$.

Now, by construction we know that $t_i^{j_*}[x] \subseteq L_{q_{k_*}}$ and $t_i^{j_*}[x] \subseteq L_{p_{j_*}}$, where x is as at the time when Test (4) is executed. However, now the language $L_{p_{j_*}}$ obtains the new element $a^i b^{x+1}$ while $L_{q_{k_*}}$ does *not*. Consequently, we have $L_{p_{j_*}} \not\subseteq L_{q_{k_*}}$. So, we have an initial segment of a text for $L_{q_{k_*}}$, i.e., $t_i^{j_*}[x]$, such that the IIM M_{i-1} , using the help-word β , cannot strong-monotonically learn $L_{q_{k_*}}$ from any text extending $t_i^{j_*}[x]$, since it has output p_{j_*} on $t_i^{j_*}[\ell]$ for some $\ell \leq x$. This is a contradiction.

Consequently, for each help-word β there is at least one language in Block i that is not learned at all by M_{i-1} or for which M_{i-1} is not strong-monotonic.

The rest is done *mutatis mutandis* as in the proof of Theorem 3 and therefore omitted. \square

Having these general results, we can also ask what happens if we allow a suitably chosen hypothesis space for \mathcal{REG} such as all DFAs. Then for all $i, j \in \mathbb{N}$ equality $L_i = L_j$ and subset $L_i \subseteq L_j$ are decidable, and thus we are in the setting described in the proof of Theorem 1. That is we have the following theorem.

Theorem 6. *Let $\mathcal{C} \subseteq \mathcal{REG}$ be arbitrarily fixed, let $d \in \mathcal{R}_{mon}$ be any function, and let $\mathcal{H} = (L_j)_{j \in \mathbb{N}}$ be any indexed hypothesis space for \mathcal{C} such that \mathcal{H} has a decidable subset problem. Then there is a computable IIM M such that the class \mathcal{C} can be strong-monotonically identified with nonconstructivity $\log d_{inv}(n)$ from text with respect to \mathcal{H} .*

3.2. Nonconstructive Learning of Recursively Enumerable Classes

Next, we turn our attention to the amount of nonconstructivity needed to learn recursively enumerable classes of recursively enumerable languages.

Theorem 7. *Let $\psi \in \mathcal{P}^2$ be any numbering. Then there is always an IIM M learning the family $(W_i^\psi)_{i \in \mathbb{N}^+}$ in the limit from text with nonconstructivity $\log n$ with respect to $(W_i^\psi)_{i \in \mathbb{N}^+}$.*

Proof. The help-word w is essentially the same as in the proof of Theorem 2, i.e., it is a bitstring of length $\log n$ which is the binary representation of ℓ , the number of pairwise different languages enumerated in $W_1^\psi, \dots, W_n^\psi$ plus n .

Let $L \in \mathcal{C} \cap \{W_1^\psi, \dots, W_n^\psi\}$ and let t be any text for L . On input any $t[m]$ and the help-word w the desired IIM M executes the following.

- (1) For all $0 < i \leq n$ enumerate W_i^ψ for m steps, that is M tries to compute $\psi_i(0), \dots, \psi_i(m)$ for at most m steps and enumerate those arguments x for which $\psi_i(x)$ turns out to be defined. Let $W_{i,m}^\psi$ be the resulting sets, $0 < i \leq n$.
- (2) For all pairs (i, j) with $0 < i, j \leq n$ check whether or not $W_{i,m}^\psi \setminus W_{j,m}^\psi \neq \emptyset$. If it is, let $d(i, j)$ be the least element in $W_{i,m}^\psi \setminus W_{j,m}^\psi$. If there is no such element, we set $d(i, j) = \infty$.
- (3) Having the numbers $d(i, j)$ the IIM M checks whether or not there are at least ℓ pairwise different languages among $W_{1,m}^\psi, \dots, W_{n,m}^\psi$. If not, then $M(t[m]) = 0$.

Otherwise, let $S = \{i \mid 0 < i \leq n, W_{i,m}^\psi \neq \emptyset\}$ and consider all sets $\tilde{S} \subseteq S$ satisfying $|\tilde{S}| = \ell$. For each such set $\tilde{S} = \{j_1, \dots, j_\ell\}$ compute the numbers $x_{j,k} =_{df} \min(W_{j,m}^\psi \Delta W_{k,m}^\psi)$ for all $j, k \in \tilde{S}$, where $j < k$ and let $s(\tilde{S})$ be the maximum of all those $x_{j,k}$. Furthermore, for each set \tilde{S} we consider the ℓ -tuple (j_1, \dots, j_ℓ) , where $j_i < j_{i+1}$, $i = 1, \dots, \ell - 1$. Using these tuples, we can order them lexicographically and then choose the first set \tilde{S} in this order for which $s(\tilde{S})$ is minimized, i.e., $s(\tilde{S}) \leq s(\hat{S})$ for all \hat{S} with $\hat{S} \subseteq S$ and $|\hat{S}| = \ell$. Let i_1, \dots, i_ℓ be the elements of this set \tilde{S} in their natural order.

Then M takes the languages $W_{i_1, m}^\psi, \dots, W_{i_\ell, m}^\psi$ into consideration. From these candidate hypotheses i_1, \dots, i_ℓ the least i is output for which $t[m]$ contains all finite $d(i, j)$, $j = i_1, \dots, i_\ell$, and $t[m]$ does not contain any of the finite $d(j, i)$, $j = i_1, \dots, i_\ell$. If there is no such i , then $M(t[m]) = 0$.

It remains to argue that M learns L in the limit from t . Note that the ℓ pairwise different languages are found in the limit, since the minimal element in the symmetric difference of the two languages tends to infinity if the two languages are equal (if any element is found at all).

So, the set of candidate hypotheses stabilizes in the limit, and by construction M then outputs the correct i as soon as the initial segment is large enough. We omit further details. \square

Note that the IIM defined in the proof of Theorem 7 even witnesses a much stronger result, i.e., it always converges to the minimum index i of the target language.

The following lower bound shows that Theorem 7 cannot be improved substantially. In order to show this result we need the following proposition.

Proposition 1. *Given a set $\mathcal{M} = \{M_1, M_2, M_3, \dots, M_m\}$ of m IIMs, one can effectively find grammars $g_0^{\mathcal{M}}, g_1^{\mathcal{M}}, \dots, g_m^{\mathcal{M}}$ such that none of M_1, M_2, \dots, M_m learns in the limit all the languages $W_{g_0^{\mathcal{M}}}, W_{g_1^{\mathcal{M}}}, \dots, W_{g_m^{\mathcal{M}}}$.*

Theorem 8. *There is a numbering $\psi \in \mathcal{P}^2$ such that no IIM M can learn the family $(W_i^\psi)_{i \in \mathbb{N}^+}$ in the limit from text with nonconstructivity $\log n - 2$ with respect to $(W_i^\psi)_{i \in \mathbb{N}^+}$.*

Proof. Using Proposition 1, the theorem can be shown as follows.

For any learner M , let M^β be the learner which behaves just like M with the advice β as additional input.

We define the numbering ψ as follows. Let φ be any fixed Gödel numbering of \mathcal{P} . We interpret every partial recursive function φ_i as an IIM M , i.e., $M_i = \varphi_i$ for every $i \in \mathbb{N}$.

The set W_j^ψ , for $2^i \leq j < 2^{i+1}$, is defined to be $W_{g_{j-2^i}^{\mathcal{M}}}$, where $\mathcal{M} = \{M_i^\beta \mid |\beta| < i\}$. Note that there are $2^i - 1$ many help-words β with $0 \leq |\beta| < i$, and we thus have the needed number 2^i of grammars (cf. Proposition 1). Thus, for any i , and any β of length at most $i - 1$, the IIM M_i^β does not learn at least one of the languages W_j^ψ , where $2^i \leq j < 2^{i+1}$.

Now, suppose by way of contradiction that there is an IIM M that learns the whole family $(W_i^\psi)_{i \in \mathbb{N}^+}$ using nonconstructivity $\log n - 2$. Then there must be an $i \in \mathbb{N}$ such that $M = M_i$.

So this IIM M_i learns the family $(W_i^\psi)_{i \in \mathbb{N}^+}$ using nonconstructivity $\log n - 2$. Then, the IIM M_i^β , for some β with length at most $i - 1$, learns all the languages W_j^ψ , for $2^i \leq j < 2^{i+1}$. A contradiction. \square

So, it remains to show Proposition 1. We shall show a bit stronger result.

Proposition 2. *Given a set $\mathcal{M} = \{M_1, M_2, M_3, \dots, M_m\}$ of m IIMs, and two disjoint finite sets S_1 and S_2 , $S_1, S_2 \subseteq \mathbb{N}$, one can effectively (in \mathcal{M}) find grammars $g_0^{\mathcal{M}, S_1, S_2}, g_1^{\mathcal{M}, S_1, S_2}, \dots, g_m^{\mathcal{M}, S_1, S_2}$ such that for some $i \leq m$, none of M_1, M_2, \dots, M_m learns in the limit all the languages $W_{g_0^{\mathcal{M}, S_1, S_2}}, W_{g_1^{\mathcal{M}, S_1, S_2}}, \dots, W_{g_m^{\mathcal{M}, S_1, S_2}}$, where we additionally have the properties that*

- (a) $S_1 \subseteq W_{g_i^{\mathcal{M}, S_1, S_2}}$, for $i \leq m$, and
- (b) $S_2 \cap W_{g_i^{\mathcal{M}, S_1, S_2}} = \emptyset$, for $i \leq m$.

Proof. We show below how to define $g_j^{\mathcal{M}, S_1, S_2}$, for every $0 \leq j \leq |\mathcal{M}|$. These grammars are generated implicitly by using the operator recursion theorem [5]. If $\mathcal{M} = \emptyset$ then we let $W_{g_0^{\mathcal{M}, S_1, S_2}} = S_1$.

Suppose $\mathcal{M} = \{M_1, M_2, \dots, M_m\}$, with $m > 0$, and let disjoint finite sets S_1, S_2 be given. The sets $W_{g_j^{\mathcal{M}, S_1, S_2}}$, for $1 \leq j \leq m$, are defined via a staging construction as follows. Note that the construction is different for each different input \mathcal{M}, S_1, S_2 .

Let σ_0 be a sequence with content S_1 . Enumerate S_1 in each of W_{h_j} , $j \leq m$.

It will be the invariant that at the beginning of stage s , $\text{content}(\sigma_s) = W_{h_j}$ enumerated before Stage s , for $1 \leq j \leq m$.

Go to Stage 0.

Stage s .

Dovetail Steps 1 and 2, until, if ever, Step 1 succeeds. If Step 1 succeeds, then stop Step 2 and go to Step 3.

Step 1. Search for an extension σ of σ_s , where $\text{content}(\sigma) \cap S_2 = \emptyset$, such that $M_{(s \bmod m)+1}(\sigma_s) \neq M_{(s \bmod m)+1}(\sigma)$.

Step 2. Let x be such that $x \notin \text{content}(\sigma_s) \cup S_2$.

Enumerate x in W_{h_m} .

Let $S'_2 = S_2 \cup \{x\}$. Let $S'_1 = \text{content}(\sigma_s)$.

Let $\mathcal{M}' = \mathcal{M} - \{M_{(s \bmod m)+1}\}$.

Let W_{h_j} , for $j < m$, follow $W_{g_j^{\mathcal{M}', S'_1, S'_2}}$ (until, if ever search in Step 1 succeeds).

Step 3. If and when Step 1 succeeds, let $X = \bigcup_{j \leq m} W_{h_j} \cup \text{content}(\sigma)$ enumerated upto now.

Enumerate X in W_{h_j} , for $j \leq m$.

Let σ_{s+1} be an extension of σ such that $\text{content}(\sigma_{s+1}) = X$.

Go to Stage $s + 1$

End *Stage s .*

We now claim that $g_j^{\mathcal{M}, S_1, S_2}$ as defined above satisfy the proposition. This is by induction on size of \mathcal{M} . Suppose the proposition properties hold for all \mathcal{M} of size less than m . Then, we show that it holds for \mathcal{M} of size m .

Fix \mathcal{M} (of size m), and S_1, S_2 and consider the following cases for the construction given above.

Case 1. There are infinitely many stages.

In this case, all $W_{h_j}, j \leq m$, are the same language given by $\text{content}(\bigcup_{s \in \mathbb{N}} \sigma_s)$, and each of the learners M_1, M_2, \dots, M_m , makes infinitely many mind changes on the text $\bigcup_{s \in \mathbb{N}} \sigma_s$.

Case 2. Stage s starts but does not finish.

In this case, clearly, $M_{(s \bmod m)+1}$ converges on all texts extending σ_s which do not contain any element from S_2 . Thus, either $M_{(s \bmod m)+1}$ does not learn W_{h_m} , (which contains x as defined in Stage s), or does not learn any of $W_{h_j}, j < m$, which do not contain x . Furthermore, by the induction hypothesis, none of the learners in \mathcal{M}' learns all the languages in W_{h_j} , where $0 \leq j < m$.

Thus, the proposition follows. \square

The situation considerably changes if we require conservative learning. In order to present this result, we need the following. A function $h: \mathbb{N} \rightarrow \mathbb{N}$ is said to be *limiting recursive* if there is a function $\tilde{h} \in \mathcal{R}^2$ such that $h(i) = \lim_{t \rightarrow \infty} \tilde{h}(i, t)$ for all $i \in \mathbb{N}$.

Theorem 9. *For every limiting recursive function h there is a recursively enumerable family $(W_i^\psi)_{i \in \mathbb{N}}$ of recursively enumerable languages such that no IIM with nonconstructivity at most h can learn $(W_i^\psi)_{i \in \mathbb{N}}$ conservatively with respect to $(W_i^\psi)_{i \in \mathbb{N}}$.*

Proof. Let the limiting recursive function $h: \mathbb{N} \rightarrow \mathbb{N}$ be arbitrarily fixed. Furthermore, we assume any fixed function $\tilde{h} \in \mathcal{R}^2$ such that $h(i) = \lim_{n \rightarrow \infty} \tilde{h}(i, n)$ for all $i \in \mathbb{N}$.

Again, we fix any Gödel numbering φ of \mathcal{P} and interpret each φ_i as an IIM, i.e., we have $M_i = \varphi_i$ for all $i \in \mathbb{N}$.

For every number $i \in \mathbb{N}$ we construct two recursively enumerable sets W_{2i}^ψ and W_{2i+1}^ψ such that for each help-word $\beta \in \{0, 1\}^*$ with $0 \leq |\beta| \leq h(2i+1)$ there will be a text such that M_i fails to learn at least one of the languages W_{2i}^ψ and W_{2i+1}^ψ or M_i is not conservative. This is done as follows.

For each number $i \in \mathbb{N}$ do the following. Set $t := 0$.

Initialize $W_{2i}^\psi = \{\langle i, 2x \rangle \mid x \in \mathbb{N}\}$ and $W_{2i+1}^\psi = \{\langle i, 2x+1 \rangle \mid x \in \mathbb{N}\}$.

Goto (1).

(1) Compute $m_t := \tilde{h}(2i+1, t)$ and let $r_t = 2^{m_t+1} - 1$.

Set $C_t = \{1, \dots, r_t\}$ and enumerate in lexicographic order all $\beta \in \{0, 1\}^*$ with $0 \leq |\beta| \leq m_t$. We refer to these elements as β_j , where $j \in C_t$.

Dovetail the execution of (α) and (2).

(α) **do forever**

$t := t + 1$;

if $\tilde{h}(2i + 1, t - 1) < \tilde{h}(2i + 1, t)$ **then**

let $m_t := \tilde{h}(2i + 1, t)$ and $r_t = 2^{m_t+1} - 1$;

Update $C_t := \{1, \dots, r_t\}$ and update (A) by including the newly found $j \in C_t$.

end

(2) Start (A).

(A) For each $j \in C_t$ try to find, by dovetailing the following search, a finite sequence σ_{β_j} of elements from $\{\langle i, x \rangle \mid x \in \mathbb{N}\}$ such that the condition $M_i(\sigma_{\beta_j}, \beta_j) \downarrow = 2i$ is verified.

Once such a sequence σ_{β_j} is found, if ever, stop (A) for the j for which the search succeeded but continue for the remaining values of j , and update

$$W_{2i}^\psi := W_{2i}^\psi \cup \text{content}(\sigma_{\beta_j}), \quad (5)$$

$$W_{2i+1}^\psi := W_{2i+1}^\psi \cup \text{content}(\sigma_{\beta_j}). \quad (6)$$

Now start (B) for the j for which the search succeeded.

(B) Try to find a finite sequence τ_{β_j} of elements from $\{\langle i, x \rangle \mid x \in \mathbb{N}\}$ such that $M_i(\sigma_{\beta_j} \diamond \tau_{\beta_j}, \beta_j) \downarrow = 2i + 1$, where $\sigma_{\beta_j} \diamond \tau_{\beta_j}$ denotes the concatenation of the finite sequences σ_{β_j} and τ_{β_j} .

If such a τ_{β_j} is found, if ever, stop (B) for the j for which the search succeeded and update

$$W_{2i}^\psi := W_{2i}^\psi \cup \text{content}(\tau_{\beta_j}), \quad (7)$$

$$W_{2i+1}^\psi := W_{2i+1}^\psi \cup \text{content}(\tau_{\beta_j}). \quad (8)$$

Furthermore, set $C_t := C_t \setminus \{j\}$. If $C_t \neq \emptyset$ continue (A) or (B) for the remaining $j \in C_t$.

If $C_t = \emptyset$ then stop (A) and (B) until, if ever, in (α) the set C_t is updated.

Now, it is easy to see that $(W_\ell^\psi)_{\ell \in \mathbb{N}}$ is a recursively enumerable family of recursively enumerable languages. It remains to show that no IIM M can learn $(W_\ell^\psi)_{\ell \in \mathbb{N}}$ conservatively with nonconstructivity h with respect to $(W_\ell^\psi)_{\ell \in \mathbb{N}}$.

Suppose the converse, i.e., there is an IIM M that learns $(W_\ell^\psi)_{\ell \in \mathbb{N}}$ conservatively with respect to $(W_\ell^\psi)_{\ell \in \mathbb{N}}$. In order to obtain a contradiction, it suffices to show that there is some $n \in \mathbb{N}$ such that for each help-word w with $0 \leq |w| \leq h(n)$ there is a language W from our family such that $W \in \{W_0^\psi, \dots, W_n^\psi\}$ but W is not learned conservatively by M .

Since M is an IIM, there must be an $i \in \mathbb{N}$ such that $M = M_i$. Furthermore, since h is limiting recursive, there must be $t_0 \in \mathbb{N}$ such that $h(2i + 1) = \tilde{h}(2i + 1, t_0 + k)$

for all $k \in \mathbb{N}$. That is, as soon as our t in (α) satisfies $t \geq t_0$ no further update of C_t is occurring in (α) . It should also be noted that we may have included more help-words β_j in our construction than actually necessary, since we only checked if $\tilde{h}(2i+1, t)$ does increase. But this does not matter as we shall see below.

We claim that for $n = 2i+1$ the desired contradiction can be obtained. We consider the sets W_{2i}^ψ and W_{2i+1}^ψ .

As said above, as soon as $t \geq t_0$ the set C_t does not change any further in the loop (α) . So, let $m = h(2i+1)$ and $r = 2^{m+1} - 1$. It suffices to consider all help-words β_1, \dots, β_r , where $0 \leq |\beta_j| \leq m$ for all $j = 1, \dots, r$.

Let any such β_j be arbitrarily fixed. Then we try in (A) in particular any finite sequence σ of elements of W_{2i}^ψ . Thus, if for all sequences σ either $M_i(\sigma, \beta_j)$ remains undefined or turns out to be defined but $M_i(\sigma, \beta_j) \neq 2i$ then M_i fails to learn W_{2i}^ψ when using the help-word β_j , since there is no other index for this language in $(W_\ell^\psi)_{\ell \in \mathbb{N}}$. Here it should be noted that $W_{2i}^\psi \neq W_{2i+1}^\psi$, since the at most finitely often occurring updates in (5), (6), (7), and (8) do only add finitely often finitely many elements to W_{2i}^ψ and W_{2i+1}^ψ , respectively.

Otherwise, the search must eventually succeed, i.e., we find a finite sequence σ_{β_j} such that $M_i(\sigma, \beta_j) \downarrow = 2i$.

Now, by construction we know that $\text{content}(\sigma_{\beta_j}) \subseteq W_{2i+1}^\psi$, too (see (6)). So, in (B) we try in particular any finite sequence τ of elements of W_{2i+1}^ψ , and by construction also each sequence $\sigma_{\beta_j} \diamond \tau$ is an initial segment of a text for W_{2i+1}^ψ . Consequently, if we never verify $M_i(\sigma_{\beta_j} \diamond \tau, \beta_j) \downarrow = 2i+1$ in (B) then M_i fails to learn W_{2i+1}^ψ from every text starting with σ_{β_j} .

Otherwise, in (B) we find a τ_{β_j} such that $M_i(\sigma_{\beta_j} \diamond \tau_{\beta_j}, \beta_j) \downarrow = 2i+1$.

Now, by construction we know that $\text{content}(\sigma_{\beta_j}) \cup \text{content}(\tau_{\beta_j}) \subseteq W_{2i+1}^\psi$ as well as $\text{content}(\sigma_{\beta_j}) \cup \text{content}(\tau_{\beta_j}) \subseteq W_{2i}^\psi$ (see (8) and (7)).

Thus, $\sigma_{\beta_j} \diamond \tau_{\beta_j}$ is an initial segment of a text for both W_{2i}^ψ and W_{2i+1}^ψ . But since

$$M_i(\sigma_{\beta_j}, \beta_j) \downarrow = 2i \neq 2i+1 = M_i(\sigma_{\beta_j} \diamond \tau_{\beta_j}, \beta_j) \downarrow ,$$

we see that $M_i(\sigma_{\beta_j}, \beta_j) \neq M_i(\sigma_{\beta_j} \diamond \tau_{\beta_j}, \beta_j)$ even though $\text{content}(\sigma_{\beta_j} \diamond \tau_{\beta_j})$ is contained in W_{2i}^ψ . This is a contradiction to M_i (with help-word β_j) being conservative.

So, for every help-word β_j , $0 \leq |\beta_j| \leq m$ the IIM M_i either does not learn W_{2i}^ψ or W_{2i+1}^ψ or it is not conservative. Thus, the theorem is shown. \square

Since $SmonTxt \subset ConsvTxt$, Theorem 9 directly allows for the following corollary.

Corollary 1. *For every limiting recursive function h there is a recursively enumerable family $(W_i^\psi)_{i \in \mathbb{N}}$ of recursively enumerable languages such that no IIM with nonconstructivity at most h can learn $(W_i^\psi)_{i \in \mathbb{N}}$ strong-monotonically with respect to $(W_i^\psi)_{i \in \mathbb{N}}$.*

4. Conclusions

We have presented a model for the inductive inference of formal languages from text that incorporates a certain amount of nonconstructivity. In our model, the amount of nonconstructivity needed to solve the learning problems considered has been used as a quantitative characterization of their difficulty.

We studied the problem of learning indexable classes under three postulates, i.e., *learning in the limit*, *conservative identification*, and *strong-monotonic inference*. As far as learning in the limit is concerned, the amount of nonconstructivity needed to learn any indexable class can be very small and there is no smallest amount that can be described in a computable way (cf. Theorem 1).

Moreover, we showed upper and lower bounds for conservative learning of indexable classes and for strong-monotonic inference roughly showing that the amount of nonconstructivity needed is $\log n$ for conservative learning and $2 \log n$ for strong-monotonic inference.

However, if we allow canonical indexed hypothesis spaces for \mathcal{REG} such that equality of languages is decidable, then the amount of nonconstructivity needed to learn \mathcal{REG} even strong-monotonically can be made very small.

Finally, we studied the problem to learn recursively enumerable classes of recursively enumerable languages. In this setting, the amount of nonconstructivity needed to learn in the limit is $\log n$, while there is not even a limiting recursive bound for the amount of nonconstructivity to learn all recursively enumerable classes of recursively enumerable languages conservatively or strong-monotonically.

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