

# Extracting Kolmogorov Complexity with Applications to Dimension Zero-One Laws

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## Abstract

We apply results on extracting randomness from independent sources to “extract” Kolmogorov complexity. For any  $\alpha, \epsilon > 0$ , given a string  $x$  with  $K(x) > \alpha|x|$ , we show how to use a constant number of advice bits to efficiently compute another string  $y$ ,  $|y| = \Omega(|x|)$ , with  $K(y) > (1 - \epsilon)|y|$ . This result holds for both unbounded and space-bounded Kolmogorov complexity.

We use the extraction procedure for space-bounded complexity to establish zero-one laws for the strong dimension of complexity classes within ESPACE. We also obtain a similar result for constructive strong dimension.

## 1 Introduction

Kolmogorov complexity quantifies the amount of randomness in an individual string. If a string  $x$  has Kolmogorov complexity  $m$ , then  $x$  is often said to contain  $m$  bits of randomness. Can we efficiently extract the Kolmogorov randomness from a string? That is, given  $x$ , is it possible to compute a string of length  $m$  that is Kolmogorov-random?

Vereshchagin and Vyugin showed that this is not possible in general [27], i.e., they showed that there is no algorithm that can extract Kolmogorov complexity. Buhrman, Fortnow, Newman and Vereshchagin [4] showed that if one allows a small amount of extra information then Kolmogorov extraction is indeed possible. More specifically, they showed there is an efficient procedure  $\mathcal{A}$  such

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\*This research was supported in part by NSF grant 0515313.

†This research was supported in part by NSF grant 0430807.

‡This research was supported in part by NSF grant 0430991.

§Work done while the author was at Iowa State University. Research supported in part by NSF grant 0430807.

that for every  $x$  with Kolmogorov complexity  $\alpha n$ , there exists a string  $a_x$ , such that  $\mathcal{A}(x, a_x)$  outputs a nearly Kolmogorov random string whose length is close to  $\alpha n$ . Moreover, the length of  $a_x$  is  $O(\log |x|)$ , and contents of  $a_x$  depend on  $x$ .

In this paper we show that we can extract Kolmogorov complexity with only *constant* bits of additional information. We give a *polynomial-time computable procedure* which takes  $x$  with an additional constant amount of advice and outputs a nearly Kolmogorov-random string whose length is linear in  $m$ . Formally, for any  $\alpha, \epsilon > 0$ , given a string  $x$  with  $K(x) > \alpha|x|$ , we show how to use a constant number of advice bits to compute another string  $y$ ,  $|y| = \Omega(|x|)$ , in polynomial-time that satisfies  $K(y) > (1 - \epsilon)|y|$ . The number of advice bits depends only on  $\alpha$  and  $\epsilon$ , but the content of the advice depends on  $x$ . This computation needs only polynomial time, and yet it extracts unbounded Kolmogorov complexity.

Our proofs use a construction of a *multi-source extractor*. Traditional extractor results [17, 28, 24, 16, 26, 20, 21, 25, 10, 23, 22, 5] show how to take a distribution with high min-entropy and some truly random bits to create a close to uniform distribution. A multi-source extractor takes several independent distributions with high min-entropy and creates a close to uniform distribution. Thus multi-source extractors eliminate the need for a truly random source. Substantial progress has been made recently in the construction of efficient multi-source extractors [2, 3, 19, 18]. In this paper we use the construction due to Barak, Impagliazzo, and Wigderson [2] for our main result on extracting Kolmogorov complexity.

To make the connection, consider the uniform distribution on the set of strings  $x$  whose Kolmogorov complexity is at most  $m$ . This distribution has min-entropy about  $m$  and  $x$  acts like a random member of this set. We can define a set of strings  $x_1, \dots, x_k$  to be independent if  $K(x_1 \dots x_k) \approx K(x_1) + \dots + K(x_k)$ . By symmetry of information this implies  $K(x_i | x_1, \dots, x_{i-1}, x_{i+1}, \dots, x_k) \approx K(x_i)$ . Suppose we are given independent Kolmogorov random strings  $x_1, \dots, x_k$ , whose Kolmogorov complexity is  $m$ . We view them as arising from  $k$  independent distributions, each with min-entropy  $m$ . We then argue that a multi-source extractor with small error can be used to output a nearly Kolmogorov random string.

To extract the randomness from a single string  $x$ , we break  $x$  into a number of substrings  $x_1, \dots, x_l$ , and view each substring  $x_i$  as coming from a different random source. Of course, these substrings may not be independently random in the Kolmogorov sense, thus we can not view these strings as coming from independent sources. A useful concept is to quantify the *dependency within*  $x$  as  $\sum_{i=1}^l K(x_i) - K(x)$ . We show that if the dependency within  $x$  is small, then the output of the multi-source extractor on its substrings is a nearly Kolmogorov random string. Another technical problem is that the randomness in  $x$  may not be nicely distributed among the substrings; for this we need to use a small (constant) number of nonuniform advice bits.

This result about extracting Kolmogorov-randomness also holds for polynomial-space bounded Kolmogorov complexity. We apply this to obtain zero-one laws for the strong dimensions of certain complexity classes. Resource-bounded dimension and strong dimension [11, 1] were developed as extensions of the classical Hausdorff and packing fractal dimensions to study the structure of complexity classes. Dimension and strong dimension both refine resource-bounded measure and are duals of each other in many ways. Strong dimension is also related to resource-bounded category [8]. In this paper we focus on strong dimension.

The strong dimension of each complexity class is a real number between zero and one inclusive. While there are examples of nonstandard complexity classes with fractional dimensions [1], we do not know of a standard complexity class with this property. Can a natural complexity class have a fractional dimension? In particular consider the class E. Determining its strong dimension within ESPACE would imply a major separation. However, we are able to use our Kolmogorov-

randomness extraction procedure to obtain a zero-one law ruling out the intermediate fractional possibility. Formally, we show that the strong dimension  $\text{Dim}(\text{E} \mid \text{ESPACE})$  is either 0 or 1. The zero-one law also holds for various other complexity classes.

Our techniques also apply in the constructive dimension setting [12]. Miller and Nies [14] asked if it is possible to compute a set of higher constructive dimension from an arbitrary set of positive constructive dimension. We answer the strong dimension variant of this question.

## 2 Preliminaries

### 2.1 Kolmogorov Complexity

Let  $M$  be a Turing machine. Let  $f : \mathbb{N} \rightarrow \mathbb{N}$ . For any  $x \in \{0, 1\}^*$ , define

$$K_M(x) = \min\{|\pi| \mid M(\pi) \text{ prints } x\}$$

and

$$KS_M^f(x) = \min\{|\pi| \mid M(\pi) \text{ prints } x \text{ using at most } f(|x|) \text{ space}\}.$$

There is a universal machine  $U$  such that for every machine  $M$ , there is some constant  $c$  such that for all  $x$ ,  $K_U(x) \leq K_M(x) + c$  and  $KS_U^{cf+c}(x) \leq KS_M^f(x) + c$  [9]. We fix such a machine  $U$  and drop the subscript, writing  $K(x)$  and  $KS^f(x)$ , which are called the (*plain*) *Kolmogorov complexity* of  $x$  and *f-bounded (plain) Kolmogorov complexity* of  $x$ . While we use plain complexity in this paper, our results also hold for prefix-free complexity.

The following definition quantifies the fraction of randomness in a string.

**Definition.** For a string  $x$ , the *rate* of  $x$  is  $\text{rate}(x) = K(x)/|x|$ . For a polynomial  $g$ , the *g-rate* of  $x$  is  $\text{rate}^g(x) = KS^g(x)/|x|$ .

We denote the uniform distribution over  $\Sigma^n$  with  $U_n$ . Two distributions  $X$  and  $Y$  over  $\Sigma^n$ , are  $\epsilon$ -close if

$$\frac{1}{2} \sum_{x \in \Sigma^n} |X(x) - Y(x)| \leq \epsilon.$$

**Definition.** Let  $X$  be a distribution over  $\Sigma^n$  and  $\text{Sup}(X)$  denotes the set  $\{x \in \Sigma^n \mid \Pr[X = x] \neq 0\}$ . The *min-entropy* of  $X$  is

$$\min_{x \in \text{Sup}(X)} \log \frac{1}{\Pr[X = x]}.$$

### 2.2 Polynomial-Space Dimension

We now review the definitions of polynomial-space dimension [11] and strong dimension [1]. For more background we refer to these papers and the survey paper [7].

Let  $s > 0$ . An *s-gale* is a function  $d : \{0, 1\}^* \rightarrow [0, \infty)$  satisfying  $2^s d(w) = d(w0) + d(w1)$  for all  $w \in \{0, 1\}^*$ .

For a language  $A$ , we write  $A \upharpoonright n$  for the first  $n$  bits of  $A$ 's characteristic sequence (according to the standard enumeration of  $\{0, 1\}^*$ ) and  $A \upharpoonright [i, j]$  for the subsequence beginning from the  $i$ th bit and ending at the  $j$ th bit. An *s-gale*  $d$  *succeeds on* a language  $A$  if  $\limsup_{n \rightarrow \infty} d(A \upharpoonright n) = \infty$  and  $d$  *succeeds strongly on*  $A$  if  $\liminf_{n \rightarrow \infty} d(A \upharpoonright n) = \infty$ . The *success set* of  $d$  is  $S^\infty[d] = \{A \mid d \text{ succeeds on } A\}$ . The *strong success set* of  $d$  is  $S_{\text{str}}^\infty[d] = \{A \mid d \text{ succeeds strongly on } A\}$ .

**Definition.** Let  $X$  be a class of languages.

1. The *pspace-dimension* of  $X$  is

$$\dim_{\text{pspace}}(X) = \inf \left\{ s \mid \begin{array}{l} \text{there is a polynomial-space computable} \\ s\text{-gale } d \text{ such that } X \subseteq S^\infty[d] \end{array} \right\}.$$

2. The *strong pspace-dimension* of  $X$  is

$$\text{Dim}_{\text{pspace}}(X) = \inf \left\{ s \mid \begin{array}{l} \text{there is a polynomial-space computable} \\ s\text{-gale } d \text{ such that } X \subseteq S_{\text{str}}^\infty[d] \end{array} \right\}.$$

For every  $X$ ,  $0 \leq \dim_{\text{pspace}}(X) \leq \text{Dim}_{\text{pspace}}(X) \leq 1$ . An important fact is that ESPACE has pspace-dimension 1, which suggests the following definitions.

**Definition.** Let  $X$  be a class of languages.

1. The *dimension of  $X$  within ESPACE* is

$$\dim(X \mid \text{ESPACE}) = \dim_{\text{pspace}}(X \cap \text{ESPACE}).$$

2. The *strong dimension of  $X$  within ESPACE* is

$$\text{Dim}(X \mid \text{ESPACE}) = \text{Dim}_{\text{pspace}}(X \cap \text{ESPACE}).$$

In this paper we will use an equivalent definition of these dimensions in terms of space-bounded Kolmogorov complexity.

**Definition.** Given a language  $L$  and a polynomial  $g$  the  *$g$ -rate of  $L$*  is

$$\text{rate}^g(L) = \liminf_{n \rightarrow \infty} \text{rate}^g(L \upharpoonright n).$$

*strong  $g$ -rate of  $L$*  is

$$\text{Rate}^g(L) = \limsup_{n \rightarrow \infty} \text{rate}^g(L \upharpoonright n).$$

**Theorem 2.1.** ([13, 6]) *Let poly denote all polynomials. For every class  $X$  of languages,*

$$\dim_{\text{pspace}}(X) = \inf_{g \in \text{poly}} \sup_{L \in X} \text{rate}^g(L).$$

and

$$\text{Dim}_{\text{pspace}}(X) = \inf_{g \in \text{poly}} \sup_{L \in X} \text{Rate}^g(L).$$

### 3 Extracting Kolmogorov Complexity

Barak, Impagliazzo, and Wigderson [2] gave an explicit multi-source extractor.

**Theorem 3.1.** ([2]) *For every constant  $0 < \sigma < 1$ , and  $c > 1$  there exist  $l = \text{poly}(1/\sigma, c)$ , a constant  $r$  and a computable function  $E : \Sigma^{\ell n} \rightarrow \Sigma^n$  such that if  $H_1, \dots, H_l$  are independent distributions over  $\Sigma^n$ , each with min entropy at least  $\sigma n$ , then  $E(H_1, \dots, H_l)$  is  $2^{-cn}$ -close to  $U_n$ , where  $U_n$  is the uniform distribution over  $\Sigma^n$ . Moreover,  $E$  runs in time  $n^r$ .*

We show that this extractor can be used to produce nearly Kolmogorov-random strings from strings with high enough complexity. The following notion of dependency is useful for quantifying the performance of the extractor.

**Definition.** Let  $x = x_1x_2 \cdots x_k$ , where each  $x_i$  is an  $n$ -bit string. The *dependency within  $x$* ,  $dep(x)$ , is defined as  $\sum_{i=1}^k K(x_i) - K(x)$ .

**Theorem 3.2.** For every  $0 < \sigma < 1$  and large enough  $n$ , there exist a constant  $l > 1$ , and a polynomial-time computable function  $E$  such that if  $x_1, x_2, \dots, x_l$  are  $n$ -bit strings with  $K(x_i) \geq \sigma n$ ,  $1 \leq i \leq l$ , then

$$K(E(x_1, \dots, x_l)) \geq n - 10l \log n - dep(x),$$

where  $x = x_1x_2 \cdots x_l$ .

*Proof.* Let  $0 < \sigma' < \sigma$ . By Theorem 3.1, there is a constant  $l$  and a polynomial-time computable multi-source extractor  $E$  such that if  $H_1, \dots, H_l$  are independent sources each with min-entropy at least  $\sigma'n$ , then  $E(H_1, \dots, H_l)$  is  $2^{-5n}$  close to  $U_n$ .

We show that this extractor also extracts Kolmogorov complexity. We prove by contradiction. Suppose the conclusion is false, i.e.,

$$K(E(x_1, \dots, x_l)) < n - 10l \log n - dep(x).$$

Let  $K(x_i) = m_i$ ,  $1 \leq i \leq l$ . Define the following sets:

$$I_i = \{y \mid y \in \Sigma^n, K(y) \leq m_i\},$$

$$Z = \{z \in \Sigma^n \mid K(z) < n - 10l \log n - dep(x)\},$$

$$Small = \{\langle y_1, \dots, y_l \rangle \mid y_i \in I_i, \text{ and } E(y_1, \dots, y_l) \in Z\}.$$

By our assumption  $\langle x_1, \dots, x_l \rangle$  belongs to  $Small$ . We use this to arrive at a contradiction regarding the Kolmogorov complexity of  $x = x_1x_2 \cdots x_l$ . We first calculate an upper bound on the size of  $Small$ .

Observe that the set  $\{xy \mid x \in \Sigma^{\sigma'n}, y \in 0^{n-\sigma'n}\}$  is a subset of each of  $I_i$ . Thus the cardinality of each of  $I_i$  is at least  $2^{\sigma'n}$ . Let  $H_i$  be the uniform distribution on  $I_i$ . Thus the min-entropy of  $H_i$  is at least  $\sigma'n$ .

Since  $H_i$ 's have min-entropy at least  $\sigma'n$ ,  $E(H_1, \dots, H_l)$  is  $2^{-5n}$ -close to  $U_n$ . Then

$$\left| P[E(H_1, \dots, H_l) \in Z] - P[U_n \in Z] \right| \leq 2^{-5n}. \quad (1)$$

Note that the cardinality of  $I_i$  is at most  $2^{m_i+1}$ , as there are at most  $2^{m_i+1}$  strings with Kolmogorov complexity at most  $m_i$ . Thus  $H_i$  places a weight of at least  $2^{-m_i-1}$  on each string from  $I_i$ . Thus  $H_1 \times \dots \times H_l$  places a weight of at least  $2^{-(m_1+\dots+m_l+l)}$  on each element of  $Small$ . Therefore,

$$P[E(H_1, \dots, H_l) \in Z] = P[(H_1, \dots, H_l) \in Small] \geq |Small| \cdot 2^{-(m_1+\dots+m_l+l)},$$

and since  $|Z| \leq 2^{n-10l \log n - dep(x)}$ , from (1) we obtain

$$|Small| < 2^{m_1+1} \times \dots \times 2^{m_l+1} \times \left( \frac{2^{n-10l \log n - dep(x)}}{2^n} + 2^{-5n} \right)$$

Without loss of generality we can take  $\text{dep}(x) < n$ , otherwise the theorem is trivially true. Thus  $2^{-5n} < 2^{-10l \log n - \text{dep}(x)}$ . Using this and the fact that  $l$  is a constant independent of  $n$ , we obtain

$$|\text{Small}| < 2^{m_1 + \dots + m_l - \text{dep}(x) - 8l \log n},$$

when  $n$  is large enough. Since  $K(x) = K(x_1) + \dots + K(x_l) - \text{dep}(x)$ ,

$$|\text{Small}| < 2^{K(x) - 8l \log n}.$$

We first observe that there is a program  $Q$  that, given the values of  $m_i$ 's,  $n$ ,  $l$ , and  $\text{dep}(x)$  as auxiliary inputs, recognizes the set  $\text{Small}$ . This program works as follows: Let  $z = z_1 \dots z_l$ , where  $|z_i| = n$ . For each program  $P_i$  of length at most  $m_i$  check whether  $P_i$  outputs  $z_i$ , by running the  $P_i$ 's in a dovetail fashion. If it is discovered that for each of  $z_i$ ,  $K(z_i) \leq m_i$ , then compute  $y = E(z_1, \dots, z_l)$ . Now verify that  $K(y)$  is at most  $n - \text{dep}(x) - 10l \log n$ . This again can be done by running programs of the length at most  $n - \text{dep}(x) - 10l \log n$  in a dovetail manner. If it is discovered that  $K(y)$  is at most  $n - \text{dep}(x) - 10l \log n$ , then accept  $z$ .

So given the values of parameters  $n$ ,  $\text{dep}(x)$ ,  $l$  and  $m_i$ 's, there is a program  $P$  that enumerates all elements of  $\text{Small}$ . Since by our assumption  $x$  belongs to  $\text{Small}$ ,  $x$  appears in this enumeration. Let  $i$  be the position of  $x$  in this enumeration. Since  $|\text{Small}|$  is at most  $2^{K(x) - 8l \log n}$ ,  $i$  can be described using  $K(x) - 8l \log n$  bits.

Thus there is a program  $P'$  based on  $P$  that outputs  $x$ . This program takes  $i$ ,  $\text{dep}(x)$ ,  $n$ ,  $m_1, \dots, m_l$ , and  $l$ , as auxiliary inputs. Since the  $m_i$ 's and  $\text{dep}(x)$  are bounded by  $n$ ,

$$\begin{aligned} K(x) &\leq K(x) - 8l \log n + 2 \log n + l \log n + O(1) \\ &\leq K(x) - 5l \log n + O(1), \end{aligned}$$

which is a contradiction. □

If  $x_1, \dots, x_l$  are independent strings with  $K(x_i) \geq \sigma n$ , then  $E(x_1, \dots, x_l)$  is a Kolmogorov random string of length  $n$ .

**Corollary 3.3.** *For every constant  $0 < \sigma < 1$ , there exists a constant  $l$ , and a polynomial-time computable function  $E$  such that if  $x_1, \dots, x_l$  are  $n$ -bit strings such  $K(x_i) \geq \sigma n$ , and  $K(x_1 x_2 \dots x_l) = \sum K(x_i) - O(\log n)$ , then  $E(x_1, \dots, x_l)$  is Kolmogorov random, i.e.,*

$$K(E(x_1, \dots, x_l)) > n - O(\log n).$$

This theorem says that given  $x \in \Sigma^{ln}$ , if each piece  $x_i$  has high enough complexity and the dependency with  $x$  is small, then we can output a string  $y$  whose Kolmogorov rate is higher than the Kolmogorov rate of  $x$ , i.e.  $y$  is relatively more random than  $x$ . What if we only knew that  $x$  has high enough complexity but knew nothing about the complexity of individual pieces or the dependency within  $x$ ? Our next theorem states that in this case also there is a procedure producing a string whose rate is higher than the rate of  $x$ . However, this procedure needs constant bits of advice.

**Theorem 3.4.** *For all real numbers  $0 < \alpha < \beta < 1$  there exist a constant  $0 < \gamma < 1$ , constants  $c, l, n_0 \geq 1$ , and a procedure  $R$  such that the following holds. For any string  $x$  with  $|x| \geq n_0$  and  $\text{rate}(x) \geq \alpha$ , there exists an advice string  $a_x$  such that*

$$\text{rate}(R(x, a_x)) \geq \min\{\text{rate}(x) + \gamma, \beta\}$$

where  $|a_x| = c$ . Moreover,  $R$  runs in polynomial time, and  $|R(x, a_x)| = \lfloor |x|/l \rfloor$ .

The number  $c$  depends only on  $\alpha, \beta$  and is independent of  $x$ . However, the contents of  $a_x$  depend on  $x$ .

*Proof.* Let  $\alpha' < \alpha$  and  $\epsilon < \min\{1 - \beta, \alpha'\}$ . Let  $\sigma = (1 - \epsilon)\alpha'$ . Using parameter  $\sigma$  in Theorem 3.2, we obtain a constant  $l > 1$  and a polynomial-time computable function  $E$  that extracts Kolmogorov complexity.

Let  $\beta' = 1 - \frac{\epsilon}{2}$ , and  $\gamma = \frac{\epsilon^2}{2l}$ . Observe that  $\gamma \leq \frac{1 - \beta'}{l}$  and  $\gamma < \frac{\alpha' - \sigma}{l}$ .

Let  $x$  have  $\text{rate}(x) = \nu \geq \alpha$ . Let  $n, k \geq 0$  such that  $|x| = ln + k$  and  $k < l$ . We strip the last  $k$  bits from  $x$  and write  $x = x_1 \cdots x_l$  where each  $|x_i| = n$ . Let  $\nu' = \text{rate}(x)$  after this change. We have  $\nu' > \nu - \gamma/2$  and  $\nu' > \alpha'$  if  $|x|$  is sufficiently large.

We consider three cases.

**Case 1.** There exists  $j$ ,  $1 \leq j \leq l$  such that  $K(x_j) < \sigma n$ .

**Case 2.** Case 1 does not hold and  $\text{dep}(x) \geq \gamma ln$ .

**Case 3.** Case 1 does not hold and  $\text{dep}(x) < \gamma ln$ .

We have two claims about Cases 1 and 2:

**Claim 3.4.1.** *Assume Case 1 holds. There exists  $i$ ,  $1 \leq i \leq l$ , such that  $\text{rate}(x_i) \geq \nu' + \gamma$ .*

*Proof of Claim 3.4.1.* Suppose not. Then for every  $i \neq j$ ,  $1 \leq i \leq l$ ,  $K(x_i) \leq (\nu' + \gamma)n$ . We can describe  $x$  by describing  $x_j$  which takes  $\sigma n$  bits, and all the  $x_i$ 's,  $i \neq j$ . Thus the total complexity of  $x$  would be at most

$$(\nu' + \gamma)(l - 1)n + \sigma n + O(\log n)$$

Since  $\gamma < \frac{\alpha' - \sigma}{l}$  and  $\alpha' < \nu'$  this quantity is less than  $\nu' ln$ . Since the rate of  $x$  is  $\nu'$ , this is a contradiction.  $\square$  *Claim 3.4.1*

**Claim 3.4.2.** *Assume Case 2 holds. There exists  $i$ ,  $1 \leq i \leq l$ ,  $\text{rate}(x_i) \geq \nu' + \gamma$ .*

*Proof of Claim 3.4.2.* By definition,

$$K(x) = \sum_{i=1}^l K(x_i) - \text{dep}(x)$$

Since  $\text{dep}(x) \geq \gamma ln$  and  $K(x) \geq \nu' ln$ ,

$$\sum_{i=1}^l K(x_i) \geq (\nu' + \gamma)ln.$$

Thus there exists  $i$  such that  $\text{rate}(x_i) \geq \nu' + \gamma$ .  $\square$  *Claim 3.4.2*

We can now describe the constant number of advice bits. The advice  $a_x$  contains the following information: which of the three cases described above holds, and

- If Case 1 holds, then from Claim 3.4.1 the index  $i$  such that  $\text{rate}(x_i) \geq \nu' + \gamma$ .
- If Case 2 holds, then from Claim 3.4.2 the index  $i$  such that  $\text{rate}(x_i) \geq \nu' + \gamma$ .

Since  $1 \leq i \leq l$ , the number of advice bits is bounded by  $O(\log l)$ . We now describe procedure  $R$ . When  $R$  takes an input  $x$ , it first examines the advice  $a_x$ . If Case 1 or Case 2 holds, then  $R$  simply outputs  $x_i$ . Otherwise, Case 3 holds, and  $R$  outputs  $E(x)$ . Since  $E$  runs in polynomial time,  $R$  runs in polynomial time.

If Case 1 or Case 2 holds, then

$$\text{rate}(R(x, a_x)) \geq \nu' + \gamma \geq \nu + \frac{\gamma}{2}.$$

If Case 3 holds, we have  $R(x, a_x) = E(x)$  and by Theorem 3.2,  $K(E(x)) \geq n - 10 \log n - \gamma \ln n$ . Since  $\gamma \leq \frac{1-\beta'}{l}$ , in this case

$$\text{rate}(R(x, a_x)) \geq \beta' - \frac{10 \log n}{n}.$$

For large enough  $n$ , this value is at least  $\beta$ . Therefore in all three cases, the rate increases by at least  $\gamma/2$  or reaches  $\beta$ .  $\square$

We now prove our main theorem.

**Theorem 3.5.** *Let  $\alpha$  and  $\beta$  be constants with  $0 < \alpha < \beta < 1$ . There exist a polynomial-time procedure  $P(\cdot, \cdot)$  and constants  $C_1, C_2, n_1$  such that for every  $x$  with  $|x| \geq n_1$  and  $\text{rate}(x) \geq \alpha$  there exists a string  $a_x$  with  $|a_x| = C_1$  such that*

$$\text{rate}(P(x, a_x)) \geq \beta$$

and  $|P(x, a_x)| \geq |x|/C_2$ .

*Proof.* We apply the procedure  $R$  from Theorem 3.4 iteratively. Each application of  $R$  outputs a string whose rate is at least  $\beta$  or is at least  $\gamma$  more than the rate of the input string. Applying  $R$  at most  $k = \lceil (\beta - \alpha)/\gamma \rceil$  times, we obtain a string whose rate is at least  $\beta$ .

Note that  $R(y, a_y)$  has output length  $|R(y, a_y)| = \lfloor |y|/l \rfloor$  and increases the rate of  $y$  if  $|y| \geq n_0$ . If we take  $n_1 = (n_0 + 1)kl$ , we ensure that in each application of  $R$  we have a string whose length is at least  $n_0$ . Each iteration of  $R$  requires  $c$  bits of advice, so the total number of advice bits needed is  $C_1 = kc$ . Thus  $C_1$  depends only on  $\alpha$  and  $\beta$ . Each application of  $R$  decreases the length by a constant fraction, so there is a constant  $C_2$  such that the length of the final outputs string is at least  $|x|/C_2$ .  $\square$

The proofs in this section also work for space-bounded Kolmogorov complexity. For this we need a space-bounded version of dependency.

**Definition.** Let  $x = x_1 x_2 \cdots x_k$  where each  $x_i$  is an  $n$ -bit string, let  $f$  and  $g$  be two space bounds. The  $(f, g)$ -bounded dependency within  $x$ ,  $\text{dep}_g^f(x)$ , is defined as  $\sum_{i=1}^k KS^g(x_i) - KS^f(x)$ .

We obtain the following version of Theorem 3.2.

**Theorem 3.6.** *For every polynomial  $g$  there exists a polynomial  $f$  such that for every  $0 < \sigma < 1$ , there exist a constant  $l > 1$ , and a polynomial-time computable function  $E$  such that if  $x_1, \dots, x_l$  are  $n$ -bit strings with  $KS^f(x_i) \geq \sigma n$ ,  $1 \leq i \leq l$ , then*

$$KS^g(E(x_1, \dots, x_l)) \geq n - 10l \log n - \text{dep}_g^f(x).$$

Similarly we obtain the following extension of Theorem 3.5.

**Theorem 3.7.** *Let  $g$  be a polynomial and let  $\alpha$  and  $\beta$  be constants with  $0 < \alpha < \beta < 1$ . There exist a polynomial  $f$ , polynomial-time procedure  $R(\cdot, \cdot)$ , and constants  $C_1, C_2, n_1$  such that for every  $x$  with  $|x| \geq n_1$  and  $\text{rate}^f(x) \geq \alpha$  there exists a string  $a_x$  with  $|a_x| = C_1$  such that*

$$\text{rate}^g(R(x, a_x)) \geq \beta$$

and  $|R(x, a_x)| \geq |x|/C_2$ .

## 4 Zero-One Laws for Complexity Classes

In this section we establish a zero-one law for the strong dimensions of certain complexity classes.

**Lemma 4.1.** *Let  $g$  be any polynomial and  $\alpha, \theta$  be rational numbers with  $0 < \alpha < \theta < 1$ . Then there is a polynomial  $f$  such that if there exists  $L \in \mathbb{E}$  with  $\text{Rate}^f(L) \geq \alpha$ , then there exists  $L' \in \mathbb{E}$  with  $\text{Rate}^g(L') \geq \theta$ .*

*Proof.* Let  $\beta$  be a real number bigger than  $\theta$  and smaller than 1 and  $f = \omega(g)$ . Pick positive integers  $C$  and  $K$  such that  $(C-1)/K < 3\alpha/4$ , and  $\frac{(C-1)\beta}{C} > \theta$ . Let  $n_1 = 1, n_{i+1} = Cn_i$ .

We now define strings  $y_1, y_2, \dots$  such that each  $y_i$  is a substring of the characteristic sequence of  $L$  or is in  $0^*$ , and  $|y_i| = (C-1)n_i/K$ . While defining these strings we will ensure that for infinitely many  $i$ ,  $\text{rate}^f(y_i) \geq \alpha/4$ .

We now define  $y_i$ . We consider three cases.

**Case 1.**  $\text{rate}^f(L \upharpoonright n_i) \geq \alpha/4$ . Divide  $L \upharpoonright n_i$  into  $K/(C-1)$  segments such that the length of each segment is  $(C-1)n_i/K$ . It is easy to see that at least for one segment the  $f$ -rate is at least  $\alpha/4$ . Define  $y_i$  to be a segment with  $\text{rate}^f(y_i) \geq \alpha/4$ .

**Case 2.** Case 1 does not hold and for every  $j$ ,  $n_i < j < n_{i+1}$ ,  $\text{rate}^f(L \upharpoonright j) < \alpha$ . In this case we punt and define  $y_i = 0^{\frac{(C-1)n_i}{K}}$ .

**Case 3.** Case 1 does not hold and there exists  $j$ ,  $n_i < j < n_{i+1}$  such that  $\text{rate}^f(L \upharpoonright j) > \alpha$ . Divide  $L \upharpoonright [n_i, n_{i+1}]$  into  $K$  segments. Since  $n_{i+1} = Cn_i$ , length of each segment is  $(C-1)n_i/K$ .

Then it is easy to show that some segment has  $f$ -rate at least  $\alpha/4$ . We define  $y_i$  to be this segment.

Since for infinitely many  $j$ ,  $\text{rate}^f(L \upharpoonright j) \geq \alpha$ , for infinitely many  $i$  either Case 1 or Case 3 holds. Thus for infinitely many  $i$ ,  $\text{rate}^f(y_i) \geq \alpha/4$ .

By Theorem 3.7, there is a procedure  $R$  with such that given a string  $x$  with  $\text{rate}^f(x) \geq \alpha/4$ , and the advice  $a_x$ ,  $\text{rate}^g(R(x, a_x)) \geq \beta$ .

Let  $w_i = R(y_i, a_{y_i})$ . Since for infinitely many  $i$ ,  $\text{rate}^f(y_i) \geq \alpha/4$ , for infinitely many  $i$ ,  $\text{rate}^g(w_i) \geq \beta$ . Also recall that  $|w_i| = |y_i|/C_2$  for an absolute constant  $C_2$ .

**Claim 4.1.1.**  $|w_{i+1}| \geq (C-1) \sum_{j=1}^i |w_j|$ .

*Proof of Claim 4.1.1.* We have

$$\sum_{j=1}^i |w_j| \leq \frac{C-1}{KC_2} \sum_{j=1}^i n_j = \frac{C-1}{KC_2} \frac{(C^i - 1)n_1}{C-1},$$

with the equality holding because  $n_{j+1} = Cn_j$ . Also,

$$|w_{i+1}| = \frac{(C-1)n_{i+1}}{KC_2} \geq \frac{(C-1)C^i n_1}{KC_2}$$

Thus

$$\frac{|w_{i+1}|}{\sum_{j=1}^i |w_j|} > (C-1).$$

□ *Claim 4.1.1*

**Claim 4.1.2.** *For infinitely many  $i$ ,  $\text{rate}^g(w_1 \dots w_i) \geq \theta$ .*

*Proof of Claim 4.1.2.* For infinitely many  $i$ ,  $rate^g(w_i) \geq \beta$ , which means  $KS^g(w_i) \geq \beta|w_i|$  and therefore

$$KS^g(w_1 \cdots w_i) \geq \beta|w_i| - O(1).$$

By Claim 4.1.1,  $|w_i| \geq (C-1)(|w_1| + \cdots + |w_{i-1}|)$ . Thus for infinitely many  $i$ ,  $rate^g(w_1 \cdots w_i) \geq \frac{(C-1)\beta}{C} - o(1) \geq \theta$ .  $\square$  Claim 4.1.2

We define  $w_1 w_2 \cdots$  to be the characteristic sequence of  $L'$ . Then by Claim 4.1.2,  $Rate^g(L') \geq \theta$ .

Next, we argue that if  $L$  is in  $E$ , then  $L'$  is in  $E/O(1)$ . Observe that  $w_i$  depends on  $y_i$  and  $a_{y_i}$ , thus each bit of  $w_i$  can be computed by knowing  $y_i$  and  $a_{y_i}$ . Recall that  $y_i$  is either a subsegment of the characteristic sequence of  $L$  or  $0^{n_i}$ . We will know  $y_i$  if we know which of the three cases mentioned above hold. This can be given as advice. Also observe that  $y_i$  is a subsequence of  $L \upharpoonright n_{i+1}$ . Also recall that  $w_i$  can be computed from  $y_i$  in time polynomial in  $|y_i|$  using constant bits of advice  $a_{y_i}$ . Since  $|w_i| = |y_i|/C_2$  for some absolute constant  $C_2$ , the running time needed to compute  $w_i$  is also polynomial in  $|w_i|$ . Since  $L$  is in  $E$ , this places  $L'$  in  $E/O(1)$ .

Finally, we observe that the advice can be removed to obtain a language in  $E$ . Let  $I$  be the set of all  $i$  such that  $rate^g(w_1 \cdots w_i) \geq \theta$ . Let  $A$  be the set of all advice strings that are used in computing  $w_i$  from  $L \upharpoonright n_{i+1}$  for  $i \in I$ . Since  $I$  is infinite and  $A$  is finite, there must be some advice string  $a \in A$  that can be used to compute infinitely many of the  $w_i$ 's. We hardcode  $a$  into the algorithm for computing  $L'$ . Call the new language we get  $L''$ . We have  $L'' \in E$ . Infinitely often,  $L''$  will be the same as  $L'$  on a  $w_i$  stretch, and it can be different elsewhere. Observe that in the proof of Claim 4.1.2 changing the strings  $w_1, \dots, w_{i-1}$  has no effect. It follows that  $Rate^g(L'') \geq \theta$ . This completes the proof of Lemma 4.1.  $\square$

**Theorem 4.2.**  $\text{Dim}(E \mid \text{ESPACE})$  is either 0 or 1.

*Proof.* Because  $E \subseteq \text{ESPACE}$ ,  $\text{Dim}(E \mid \text{ESPACE}) = \text{Dim}_{\text{pspace}}(E)$ . We will show that if  $\text{Dim}_{\text{pspace}}(E) > 0$ , then  $\text{Dim}_{\text{pspace}}(E) = 1$ . For this, it suffices to show that for every polynomial  $g$  and real number  $0 < \theta < 1$ , there is a language  $L'$  in  $E$  with  $Rate^g(L') \geq \theta$ . By Theorem 2.1, this will show that the strong pspace-dimension of  $E$  is 1.

The assumption states that the strong pspace-dimension of  $E$  is greater than 0. If the strong pspace-dimension of  $E$  is actually one, then we are done. If not, let  $\alpha$  be a positive rational number that is less than  $\text{Dim}_{\text{pspace}}(E)$ . By Theorem 2.1, for every polynomial  $f$ , there exists a language  $L \in E$  with  $Rate^f(L) \geq \alpha$ .

By Lemma 4.1, from such a language  $L$  we obtain a language  $L'$  in  $E$  with  $Rate^g(L') \geq \theta$ . Thus the strong pspace-dimension of  $E$  is 1.  $\square$

The zero-one law in Theorem 4.2 also holds for many other complexity classes.

**Theorem 4.3.** Let  $\mathcal{C}$  be a class that is closed under exponential-time truth-table reductions. Then  $\text{Dim}(\mathcal{C} \mid \text{ESPACE})$  is either 0 or 1.

Therefore additional examples of classes the zero-one law holds for include  $\text{NE} \cap \text{coNE}$ ,  $\text{BPE}$ , and  $\text{E}^{\text{NP}}$ .

**Remark.** Theorem 4.2 concerns strong dimension. For dimension, the situation is considerably more complicated. With our techniques we can prove that if  $\text{dim}_{\text{pspace}}(E) > 0$ , then  $\text{dim}_{\text{pspace}}(E/O(1)) \geq 1/2$ . It appears that a different method is needed to eliminate the advice or increase the dimension past  $1/2$ .

## 5 Increasing Constructive Strong Dimension

Miller and Nies [14] asked if every set of positive constructive dimension computes (by way of a Turing reduction) a set of higher constructive dimension. Our techniques yield a positive answer for the variant of this question using strong dimension instead of dimension. For a set  $S$ , the constructive strong dimension [1] of  $S$  is defined by

$$\text{Dim}(S) = \limsup_{n \rightarrow \infty} \frac{K(S \upharpoonright n)}{n}.$$

**Theorem 5.1.** *If  $\text{Dim}(S) > 0$ , then for every  $\epsilon > 0$ , there exists  $R \leq_T S$  such that  $\text{Dim}(R) > 1 - \epsilon$ .*

The proof of Theorem 5.1 is the same as Lemma 4.1, except instead of Theorem 3.7 we use Theorem 3.5. The reduction we obtain is actually an exponential-time truth-table reduction, so in particular it is a weak truth-table reduction. In contrast, Nies and Reimann [15] showed that this is sometimes impossible for constructive dimension: there exists  $S$  with  $\text{dim}(S) > 0$  such that every set which weak truth-table reduces to  $S$  has  $\text{dim}(R) \leq \text{dim}(S)$ .

## Acknowledgments

We thank Xiaoyang Gu and Philippe Moser for several helpful discussions.

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