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We study the nonuniform complexity of problems decidable in exponential space (ESPACE). Lutz showed that almost every problem in ESPACE has circuit-size complexity $\Omega(2^n/n)$. We investigate the more restrictive model of disjunctive normal form (DNF) complexity. We prove that almost every problem in ESPACE has DNF complexity $\Theta(\frac{2^n}{\lg n \lg \lg n})$. Every pspace-random sequence must have DNF complexity $\Theta(\frac{2^n}{\lg n \lg \lg n})$. 
POLYNOMIAL-SPACE RANDOMNESS AND DNF COMPLEXITY

by

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To my mother Adetoun Sekoni
Contents

Acknowledgments v

Chapter 1 Introduction 1
  1.1 Preliminaries .................................................. 2
  1.2 Boolean Functions ........................................... 3
    1.2.1 Hypercube ............................................. 4
  1.3 Resource-Bounded Measure .................................. 4
  1.4 Space-Bounded Kolmogorov Complexity .................... 5

Chapter 2 Resource-Bounded Measure of Nonuniform DNF complexity of Languages 6

References 14
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Chapter 1

Introduction

The disjunctive normal form (DNF) complexity of Boolean functions has been extensively studied with many fundamental results obtained. Of particular importance to us is the Korshunov-Kuznetsov theorem [1]. It states that the optimal DNF size for a random Boolean function $f : \{0, 1\}^n \rightarrow \{0, 1\}$ is

$$(K + o(1))(2^{n/\lg n \lg \lg n}),$$

where $1 \leq K \leq 1.54169$.

Nigmatullin showed that for almost all Boolean functions $f : \{0, 1\}^n \rightarrow \{0, 1\}$ and $\epsilon > 0$

$$(1 - \epsilon)\bar{\ell}(n) \leq \ell(f) \leq (1 + \epsilon)\bar{\ell}(n).$$

Here $\ell(f)$ is the number of terms in the smallest DNF that computes $f$ and $\bar{\ell}(n)$ is the average of $\ell(f)$ when $f$ is uniformly selected from the set of all Boolean functions from $\{0, 1\}^n$ to $\{0, 1\}$. We extend the knowledge of DNF complexity to the resource-bounded measure framework. Similar work has been carried out by Lutz [2], in the more general setting of circuit-size complexity.

Circuit-size complexity has also been extensively studied. Shannon [3] showed that every Boolean function $f : \{0, 1\}^n \rightarrow \{0, 1\}$ can be computed by a circuit of size $O(2^n/n)$. Lupanov [4] showed that every Boolean function $f : \{0, 1\}^n \rightarrow \{0, 1\}$ is computed by a circuit of size $(2^n/n)(1 + O(1/\sqrt{n}))$. Using resource-bounded measure Lutz [2] showed that almost every language in exponential space has circuit-size complexity $\Omega(2^n/n)$. We show
that the DNF complexity of almost every language is within a constant factor of the average DNF complexity of a random Boolean function. We prove that all pspace-random languages have DNF complexity \( \Theta\left(\frac{2^n}{\lg n \lg \lg n}\right) \).

Using infinite binary strings to represent languages, we prove our main result by establishing a lower bound on space-bounded Kolmogorov complexity of languages with nonuniform DNF complexity \( \Theta\left(\frac{2^n}{\lg n \lg \lg n}\right) \). A language \( L \) has nonuniform DNF complexity \( NDC : \mathbb{N} \rightarrow \mathbb{R} \), if for all but finite \( n \) we can construct a DNF formula of size \( \leq NDC(n) \) such that \( s \in L \cap \{0, 1\}^n \iff \gamma(s) = 1 \).

A major component of our proof is bounding the probability that the DNF complexity of a random function deviates from \( \Theta\left(\frac{2^n}{\lg n \lg \lg n}\right) \). We do this by considering two cases for some parameter \( k \in \mathbb{N} \). The first case is that a random function has a subcube of dimension \( k \) and the second case is that it doesn’t have any subcube of dimension \( k \) and has size \( \Theta\left(\frac{2^n}{\lg n \lg \lg n}\right) \). Both probabilities are inversely related so we need to choose \( k \) in such a way that both events have a “small enough” probability. For instance in the proof of R. G. Nigmatullin’s theorem in [5], \( \lceil 2 \lg n \rceil \) is chosen for \( k \). But this value doesn’t work for our purposes and leads to a probability bound on the first case that is too big. We show that choosing \( k \) to be \( \epsilon n \) for small enough \( \epsilon \) leads to a better probability for the first case and doesn’t make the second case too bad.

1.1 Preliminaries

The set of all binary strings is \( \{0, 1\}^* \). The length of a string \( x \in \{0, 1\}^* \) is \( |x| \). The empty string is denoted as \( \lambda \). For \( n \in \mathbb{N} \), \( \{0, 1\}^n \) is the set of strings of length \( n \) and \( \{0, 1\}^{\leq n} \) is the set of strings of length at most \( n \). We write \( s_0 = \lambda, s_1 = 0, s_2 = 1, s_3 = 00, ... \) for the standard lexicographic enumeration of \( \{0, 1\}^* \).

A language is a subset \( L \subseteq \{0, 1\}^* \). We write \( L^{=n} = L \cap \{0, 1\}^n \) and \( L^{\leq n} = L \cap \{0, 1\}^{\leq n} \). Associated with every language \( L \) is its characteristic sequence \( \chi_L \in \{0, 1\}^\infty \). It is defined as

\[ \chi_L[i] = 1 \iff s_i \in L \text{ for } i \in \mathbb{N}, \]
where $\chi_{L}[i]$ is the $i^{th}$ bit of $\chi_{L}$. We can also index $\chi_{L}$ with strings as in $\chi_{L}[s_{i}] = \chi_{L}[i]$ for $i \in \mathbb{N}$. $\chi_{L}[i,j]$ denotes the $i^{th}$ through $j^{th}$ bits of $\chi_{F}$ inclusive while $\chi_{L_{=n}}$ denotes $\chi_{L}[2^{n} - 1, 2^{n+1} - 2]$, i.e. the substring of $\chi_{L}$ corresponding to the strings in $L_{=n}$.

We say that a statement $P(n)$ holds almost everywhere if it is true for all but finitely many $n \in \mathbb{N}$. It is denoted by $P(n)$ $a.e.$ Similarly we write $P(n)$ $i.o.$ to say that $P(n)$ holds for infinitely many $n$.

### 1.2 Boolean Functions

A Boolean function is any $f : \{0, 1\}^{n} \rightarrow \{0, 1\}$. Associated with any Boolean function is its characteristic sequence $\chi_{f} \in \{0, 1\}^{2^{n}}$ defined as

$$\chi_{f}(w) = 1 \iff f(w) = 1 \text{ for } w \in \{0, 1\}^{n},$$

equivalently

$$\chi_{f} = f(s_{2^{n}-1})f(s_{2^{n}}) \ldots f(s_{2^{n+1}-2})$$

Besides its characteristic sequence a Boolean function can be represented in other ways. In this paper we also represent them in disjunctive normal form (DNF). A DNF representation is the disjunction (logical OR) of zero or more terms; a term is the conjunction (logical AND) of zero or more literals; a literal is either a Boolean variable or its negation (logical NOT). Given a Boolean function $f$ we use $D_{f}$ to denote a DNF representation of $f$ and $\text{size}(D_{f})$ for the number of terms in $D_{f}$. We write $\ell(f)$ or $\ell(\chi_{f})$ for

$$\min\{\text{size}(D) | D \text{ is a DNF representation of } f\},$$

$\bar{\ell}(n)$ for

$$E[\ell(f)].$$

In the second definition the expectation is taken over the random uniform selection from the set of all Boolean functions from $\{0, 1\}^{n}$ to $\{0, 1\}$. Thus $\bar{\ell}(n)$ is the average DNF complexity of a random Boolean function on $n$ variables.
The DNF-size complexity of a language \( L \) is the function
\[
DS_L : \mathbb{N} \rightarrow \mathbb{N}
\]
\[
DS_L(n) = \ell(\chi_{L\preceq n}).
\]
I.e. \( DS_L(n) \) is the size of the smallest DNF whose characteristic sequence is \( \chi_{L\preceq n} \).
Similarly, we can define the DNF complexity \( DS_x \) for \( x \in \{0, 1\}^\infty \)
\[
DS_x : \mathbb{N} \rightarrow \mathbb{N}
\]
\[
DS_x(n) = DS_L(n),
\]
where \( L \) is the language whose characteristic sequence is \( x \).

1.2.1 Hypercube

An \( n \)-cube is an undirected graph whose vertex set is \( \{0, 1\}^n \) and has an edge between any two vertices that differ in one bit position. Sometimes we view a Boolean function \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) as a covering the \( n \)-cube. We say \( f \) covers some vertex set \( S \) of an \( n \)-cube if \( S = f^{-1}(1) \). Given an \( n \)-cube we refer to its subgraphs that are also \( k \)-cubes for \( k \in \{0, 1, \cdots, n\} \) as subcubes. It is easy to see that every subcube corresponds to some term, i.e it is covered by a term.

1.3 Resource-Bounded Measure

A set \( X \subseteq \{0, 1\}^\infty \) has pspace measure zero, denoted by \( \mu_{\text{pspace}}(X) = 0 \), if there is a pspace-computable martingale that attains arbitrarily large values on it. \( X \) has measure one if it is the complement of a measure zero set. Similarly we say that \( X \) has measure zero in ESPACE, denoted by \( \mu(X|\text{ESPACE}) = 0 \), if there is a pspace-computable martingale that attains arbitrarily large values on \( X \cap \text{ESPACE} \). \( X \) has measure one in ESPACE if its complement has measure zero in ESPACE. We say \( x \in \{0, 1\}^\infty \) is pspace-random if there is no pspace-computable martingale that attains arbitrarily large values on it. We do not go into the full definition of pspace measure as Theorem 1.1 suffices for our purpose. For a review of resource-bounded measure see [2].
1.4 Space-Bounded Kolmogorov Complexity

Given a Turing machine $M$ and “program” $\pi \in \{0,1\}^*$ for $M$ we say that “$M(\pi,n) = w$ in $\leq s$ space” if $M$, on input $(\pi,n)$, outputs the string $w \in \{0,1\}^*$ and halts without using more than $s$ cells of workspace. We are only interested in the situation where the output has the from $\chi_{L_n}$, i.e., the characteristic sequence of $L_n$, for some language $L$.

Given a Turing machine $M$, a space bound $s : \mathbb{N} \rightarrow \mathbb{N}$, a language $L$, and a natural number $n$, the $s(n)$-space-bounded Kolmogorov complexity of $L = n$ relative to $M$ is

$$K S_M^{s(n)}(L = n) = \min\{|\pi| \mid M(\pi,n) = \chi_{L_n} \text{in } \leq s(n) \text{ space}\}$$

Intuitively $K S_M^{s(n)}(L = n)$ is the length of the shortest program that causes the Turing machine $M$ to halt with output $\chi_{L_n}$ in $\leq s(n)$ space. Using standard efficient simulation techniques we can show that there is a universal machine $U$ such that for each machine $M$ there is a constant $c$ such that for all $s, L$ and $n$, we have

$$K S_U^{c,s(n)+c}(L = n) \leq K S_M^{s(n)}(L = n) + c.$$  

Henceforth we fix such a Turing machine $U$ and omit it from the notation.

**Theorem 1.1.** Lutz [6] Let $c \in \mathbb{N}$ and $\epsilon > 0$.

If

$$X = \{L \subseteq \{0,1\}^* \mid K S_{2^{cn}}(L = n) > 2^n - 2^m \text{ a.e.}\},$$

then $\mu_{\text{pspace}}(X) = \mu(X|\text{ESPACE}) = 1$. 

Chapter 2

Resource-Bounded Measure of Nonuniform DNF complexity of Languages

In this section we prove our main result and its corollary:

**Theorem 2.1.** The set of all \( x \in \{0,1\}^\infty \) such that \( DS_x(n) = \Theta\left(\frac{2^n}{\lg n \lg \lg n}\right) \) has pspace-measure 1 and measure 1 in ESPACE.

**Corollary 2.2.** Every pspace-random language has DNF complexity \( \Theta\left(\frac{2^n}{\lg n \lg \lg n}\right) \)

We treat \( n/k \) where \( n, k \in \mathbb{N} \) as integers. We don’t include floors or ceilings because they don’t affect our results but make things messy. We make use of the following bound

\[
\binom{n}{n/k} \leq 2^n H(1/k), \text{ where } H(\alpha) = \alpha \lg \frac{1}{\alpha} + (1 - \alpha) \lg \frac{1}{1 - \alpha}.
\]

Analogous to \( \ell(f) \) and \( \overline{\ell}(n) \) we define \( \ell_k^+(f) \) and \( \overline{\ell}_k^+(n) \), where \( k \) is a positive integer. \( \ell_k^+(f) \) denotes the size of the smallest DNF computing \( f \) in which no term has as few as \( \frac{k-1}{k} n \) literals. \( \overline{\ell}_k^+(n) \) denotes the expectation of \( \ell_k^+(f) \) over the uniform random distribution on Boolean functions from \( \{0,1\}^n \) to \( \{0,1\} \).

What follows are six claims that build up to our main result. The first claim establishes an upper bound on the probability that the augmented DNF size \( \overline{\ell}_k^+(n) \) of a random Boolean function \( f \) deviates from its average \( \overline{\ell}_k^+(n) \).
The proof of the following claim is based on the proof of Theorem 2.1 in [5].

Claim 2.1. $Pr[|\ell_k^+(f) - \ell_k^-(n)| > 2^{2^n}] \leq 2 \exp(-2^{n/6})$ for sufficiently large $k$.

Proof. A sequence $X_0, X_1, \cdots$ of random variables is a martingale with respect to the sequence $Z_0, Z_1, \cdots$ if, for all $n \geq 0$, the following conditions hold:

- $X_n$ is a function of $Z_0, Z_1, \cdots, Z_n$;
- $E[|X_n|] < \infty$;
- $E[X_{n+1}|Z_0, \cdots, Z_n] = X_n$.

A Doob martingale is a martingale constructed in the following way. Let $Z_0, Z_1, \cdots, Z_n$ be a sequence of random variables, and let $Y$ be a random variable with $E[|Y|] < \infty$. Then

$$X_i = E[Y|Z_0, \cdots, Z_i], \quad i = 0, 1, \cdots, n$$

gives a martingale with respect to $Z_0, Z_1, \cdots, Z_n$. [7]

Azuma’s inequality [8] states that if $0 = X_0, X_1, \cdots, X_N$ is a martingale with respect to the sequence $Z_0, Z_1, \cdots, Z_N$ and if $|X_{i+1} - X_i| \leq C \neq 0$ for $0 \leq i \leq N - 1$, then

$$Pr[|X_N| > \lambda] \leq 2 \exp\left(-\frac{\lambda^2}{2C^2 N}\right).$$

We now construct a Doob martingale as follows. Let $N = 2^n$, and let $s_1, s_2, \cdots, s_{2^n}$ be the strings of $\{0, 1\}^n$, in some arbitrary order. Let $X_i = Y_i - \ell_k^-(n)$ where $Y_i$ is the expectation of $\ell_k^+(f)$ conditioned on the values of $f$ on the strings $s_1, s_2, \cdots, s_i$. It is easy to see that $X_0 = 0$ and $X_N = \ell_k^+(f) - \ell_k^-(n)$. We choose $C$ to be the maximum of $|\ell_k^+(f) - \ell_k^-(g)|$ over all Boolean functions $f$ and $g$ that differ only on a single string in $\{0, 1\}^n$. If we flip the bit of a Boolean function on some string from a 0 to 1, then the resulting function has at most one more term in its shortest DNF formula than the original formula. On the other hand if we flip from 1 to 0 then we may have to add more terms. For sufficiently large $k$ the changed string can lie in at most $\binom{n}{(n/k)-1}$ terms in any shortest disjunctive normal form in which no term has as few as $k-1/k$ literals. Each of these terms can be replaced by $(n/k) - 1$ terms that cover all the strings of the original term with the exception of the string whose value

7
was changed to 0. Thus the number of terms added is at most $\frac{n-k}{k} \binom{n}{(n/k)} \leq 2^{nH(1/k)+\log \frac{n-k}{k}}$. Now we apply Azuma’s inequality with $N = 2^n$, $\lambda = 2^{2n/3}$ and $C = 2^{n/12}$ by choosing $k$ sufficiently large. We get

$$Pr[|\ell_k^+(f) - \ell_k^+(n)| > 2^{\frac{n}{2}}] \leq 2 \exp(-2^{n/6}), \text{ for sufficiently large } k.$$ 

The following claim shows that substituting DNF size $\ell(f)$ for the augmented DNF size $\ell_k^+(f)$ doesn’t significantly change the probability bound of the previous claim.

**Claim 2.2.** For sufficiently large $k$ and $\epsilon(k) > 0$,

$$Pr[|\ell(f) - \ell_k^+(n)| > 2^{\frac{n}{2}}] \leq 2^{-2^{(k)n}} \text{ a.e.}$$

**Proof.** Since $\ell_k^+(f) = \ell(f)$, unless $f$ covers a subcube of dimension $n/k$ we see that

$$Pr[\ell_k^+(f) \neq \ell(f)] \leq Pr[f \text{ covers a subcube of dimension } n/k].$$

The probability that a random Boolean function has a particular subcube of dimensions $n/k$ is $2^{-2^{n/k}}$. There are $\binom{n}{n/k} 2^{2\frac{1}{k}n}$ subcubes of dimension $n/k$. By the union bound the probability that a random Boolean function has a subcube of dimension $n/k$ is at most

$$2^{-2^{n/k}} \binom{n}{n/k} 2^{2\frac{1}{k}n} \leq 2^{-2^{n/k} + n(H(1/k) + \frac{1}{k})}.$$

Hence

$$Pr[|\ell(f) - \ell_k^+(n)| > 2^{\frac{n}{2}}] = Pr[|\ell(f) - \ell_k^+(n)| > 2^{\frac{n}{2}} \land \ell(f) = \ell_k^+(f)] +$$

$$Pr[|\ell(f) - \ell_k^+(n)| > 2^{\frac{n}{2}} \land \ell(f) \neq \ell_k^+(f)]$$

$$= Pr[|\ell_k^+(f) - \ell_k^+(n)| > 2^{\frac{n}{2}} \land \ell(f) = \ell_k^+(f)] +$$

$$Pr[|\ell(f) - \ell_k^+(n)| > 2^{\frac{n}{2}} \land \ell(f) \neq \ell_k^+(f)]$$

$$\leq Pr[|\ell_k^+(f) - \ell_k^+(n)| > 2^{\frac{n}{2}}] + Pr[\ell(f) \neq \ell_k^+(f)]$$

$$\leq 2 \exp(-2^{n/6}) + 2^{-2^{n/k} + n} \text{ a.e. for sufficiently large } k$$

$$\leq 2^{-2^{(k)n}} \text{ a.e. for sufficiently large } k \text{ and } \epsilon(k) = \frac{1}{2k}.$$
From here onwards we set \( k \) to \( k_0 \) and \( \epsilon \) to \( \epsilon(k_0) \) such that

\[
Pr[|\ell(f) - \ell^+(k_0)(n)| > 2^{\frac{2}{3}n}] \leq 2^{-2^n} \text{ a.e.}
\]

Using the previous probability bound, the next claim establishes an upper bound on the number of Boolean functions \( f : \{0,1\} \rightarrow \{0,1\} \) that deviate a “little” from \( \ell^+(k_0)(n) \). This bound will be used in computing an upper bound on the space-bounded Kolmogorov complexity of

\[
\{x \in \{0,1\}^\infty \mid |DS_x(n) - \ell^+(k_0)(n)| > 2^{\frac{2}{3}n} \text{ i.o.}\}
\]

in Claim 2.5.

**Claim 2.3.** The number of Boolean functions \( f : \{0,1\}^n \rightarrow \{0,1\} \) such that \( |\ell(f) - \ell^+(k_0)(n)| > 2^{\frac{2}{3}n} \) is at most \( 2^{2^n - 2^n} \) for sufficiently large \( n \).

**Proof.** This follows directly from Claim 2.1. Since the probability distribution we used on the \( 2^n \) Boolean functions is the uniform distribution, there are at most \( 2^{2^n - 2^n} \) such functions. \( \square \)

Now we present two algorithms that will be used in Claim 2.5 to bound the space-bounded Kolmogorov complexity of the languages in

\[
\{x \in \{0,1\}^\infty \mid |DS_x(n) - \ell^+(k_0)(n)| > 2^{\frac{2}{3}n} \text{ i.o.}\}
\]

Consider the following algorithms, the first computes \( \ell^+(k_0)(n) \) and the second takes in two non-negative integer arguments \( i, n \) and outputs \( \chi_f \). Here \( \chi_f \) is the characteristic string of the \( i^{th} \) Boolean function \( f : \{0,1\}^n \rightarrow \{0,1\} \) with size \( \ell(f) \) such that \( |\ell(f) - \ell^+(k_0)(n)| \leq 2^{\frac{2}{3}n} \).
Algorithm 1.

1: Input: \( n \)
2: Output: \( \ell_{k_0}^+(n) \)
3: total = 0
4: for all Boolean function \( f : \{0, 1\}^n \rightarrow \{0, 1\} \) do
5: \( \quad \) min = \( 2^n \)
6: for all DNF formulae \( D \) with no subcube of dimension \( n/k \) and size \( \leq 2^n \) do
7: \( \quad \) if \( D \) computes \( f \) then
8: \( \quad \quad \) min = minimum\{min, size(D)\}
9: \( \quad \) end if
10: end for
11: total = total + min
12: end for
13: \( \ell_{k_0}^+(n) = \frac{\text{total}}{2^n} \)
14: return \( \ell_{k_0}^+(n) \)
Algorithm 2.

1: Input: $i, n$
2: Output: $\chi_f$ where $f$ is the $i^{th}$ Boolean function with $n$ arguments such that
   \[ |\ell(f) - \ell_{k_0}^+(n)| \leq 2^\frac{2}{3} n \]
3: compute $\ell_{k_0}^+(n)$
4: count = 0
5: for all Boolean function $f : \{0,1\}^n \rightarrow \{0,1\}$ do
6:     min = $2^n$
7:     for all DNF formula $D$ with size $\leq 2^n$ do
8:         if $D$ computes $f$ then
9:             min = minimum\{min, size($D$)\}
10:        end if
11:     end for
12:     if $|\min - \ell_{k_0}^+(n)| > 2^\frac{2}{3} n$ then
13:         count = count + 1
14:     end if
15:     if count = $i$ then
16:         return $\chi_f$
17:     end if
18: end for

Claim 2.4. Algorithm 1 and Algorithm 2 use space polynomial in $2^n$.

Proof. Because of the similarity of both algorithms we’ll only consider algorithm 1. The for-loop of line 4 iterates for exactly $2^{2n}$ times thus we need $2^n$ bits to implement its counter. The inner for-loop of line 6 iterates for at most $2^{2n}$ times since there are $3^n$ possible terms which occurs at most once in any DNF formula. Therefore this loop requires at most $2^{n\log_2 3}$ bits to implement its counter. The predicate $D$ computes $f$ of line 8 can also be easily
implement in space polynomial in $2^n$. The variables $total, \text{min}, \ell_{k_0}^+(n)$ are bounded above by $2^n + 2^2$ and thus each uses at most $2^n + n$ bits.

\begin{claim}
Let $X = \{x \in \{0, 1\}^\infty \mid |DS_x(n) - \ell_{k_0}^+(n)| > 2^\frac{3}{2}n\ i.o.\}$. Then there is a constant $c_0 > 0$ such that for all $x \in X$,
\[ KS^{2^n}(L=n) \leq 2^n - 2^{cn} + c_0 \ i.o., \text{ where } \chi_L = x. \]
\end{claim}

\textbf{Proof.} Given $x \in X$ let $f$ be the Boolean function associated with the characteristic sequence $x[length n]$ and $L$ be the language whose characteristic sequence is $x$. Claim 2.3 tells us that for sufficiently large $n$ there are at most $2^n - 2^n$ Boolean functions $f : \{0, 1\}^n \longrightarrow \{0, 1\}$ such that $|\ell(f) - \ell_{k_0}^+(n)| > 2^\frac{3}{2}n$. Since by definition $\ell(f) = DS_x(n)$, there exists a positive $i \leq 2^n - 2^n$ such that Algorithm 2 outputs $x[length n]$. Since we need at most $2^n - 2^n$ bits to represent $i$ and Algorithm 2 uses at most $2^cn$ space, for some constant $c$, the definition of space-bounded Kolmogorov complexity implies that $KS^{2^n}(L=n) \leq 2^n - 2^n$ i.o. here $c_0$ is the length of the string used to encode a universal Turing machine.

The following claim is a minor extension of results proved by Pippenger [5]. It shows that $\ell_k^+(n) \approx \ell(n) = \Theta(\frac{2^n}{\lg n \lg \lg n})$.

\begin{claim}
$\ell_k^+(n) = \Theta(\frac{2^n}{\lg n \lg \lg n})$.
\end{claim}

\textbf{Proof.}
\begin{align*}
(1 + o(1)) \frac{2^n}{\lg n \lg \lg n} & \leq \ell(n) \\
\ell(n) & \leq (1 + o(1)) \frac{2^{n+1}}{\lg n \lg \lg n}
\end{align*}

(2.1) and (2.2) are due to Kuznetsov and Pippenger respectively [5]. We make use of the following observations:

$\ell_k^+(f) = \ell(f)$, unless $f$ covers a subcube of dimension $n/k$.

$\ell(f) \leq \ell_k^+(f)$

$\ell_k^+(f) - \ell(f) \leq |\{x \in \{0, 1\}^n \mid x \text{ is contained in a subcube of dimension } n/k \text{ covered by } f\}|$

$E[\ell_k^+(f) - \ell(f)] \leq \binom{n}{n/k} 2^{\frac{k-1}{k}n} 2^{-2^{k/n}2^n/k} \leq 2^{-2^{n/k} + (H(1/k) + 1)n}$

12
The last inequality follows because there are \( \binom{n}{n/k} 2^{\frac{k-1}{k}n} \) subcubes of dimension \( n/k \). Each dimension \( n/k \) cube is included with probability \( 2^{-2n/k} \) and contains \( 2^{n/k} \) points. By the linearity of expectations we also see that \( \ell(n/k) \leq \ell(n) + 2^{-2n/k} + (H(1/k)+1) \). The second observation implies that \( \ell(n/k) \geq \ell(n) \). Thus \( \ell(n/k) \) is arbitrarily close to \( \ell(n) \). Therefore

\[
\ell(n/k) = \Theta\left(\frac{2^n}{\log n \log \log n}\right).
\]

We now bring everything together using Claim 2.5 & 2.6 and Theorem 1.1 to show that \( \{x \in \{0, 1\}^\infty | DS_x = \Theta\left(\frac{2^n}{\log n \log \log n}\right)\} \) is a measure 1 set by showing that it is a superset of a measure 1 set.

**Theorem 2.1.** The set of all \( x \in \{0, 1\}^\infty \) such that \( DS_x(n) = \Theta\left(\frac{2^n}{\log n \log \log n}\right) \) a.e. has pspace-measure 1 and measure 1 in ESPACE.

**Proof.** It follows from claim 2.5 that

\[
X = \{x \in \{0, 1\}^\infty | |DS_x(n) - \ell(n/k)| > 2^{2n/3} \text{ i.o.}\} \\
\subseteq \{x \in \{0, 1\}^\infty | (KS^{2^c}(L_n) \leq 2^n - 2^{c_0} + c_0 \text{ i.o.}) \land (\chi_L = x)\} \\
\implies \\
X^c = \{x \in \{0, 1\}^\infty | |DS_x(n) - \ell(n/k)| \leq 2^{2n/3} \text{ a.e.}\} \\
\supseteq \{x \in \{0, 1\}^\infty | (KS^{2^c}(L_n) > 2^n - 2^{c_0} + c_0 \text{ a.e.}) \land (\chi_L = x)\}
\]

Theorem 1.1 tells us that

\[
\{x \in \{0, 1\}^\infty | (KS^{2^c}(L_n) > 2^n - 2^{c_0} + c_0 \text{ a.e.}) \land (\chi_L = x)\}
\]

has pspace-measure 1 and hence measure one in ESPACE. Since \( X^c \) is a superset of a pspace-measure 1 set, it follows that \( \mu_{\text{pspace}}(X^c) = \mu(X^c|\text{ESPACE}) = 1 \).

Similarly since \( Y = \{x \in \{0, 1\}^\infty | DS_x = \Theta\left(\frac{2^n}{\log n \log \log n}\right)\} \) is a superset of \( X^c \). Therefore, it follows that

\[
\mu_{\text{pspace}}(Y) = \mu(Y|\text{ESPACE}) = 1.
\]
References


