# A Comparison of Two Approaches to Pseudorandomness

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**Abstract**

The concept of pseudorandomness plays an important role in cryptography. In this note we contrast the notions of complexity-theoretic pseudorandom strings (from algorithmic information theory) and pseudorandom strings (from cryptography). For example, we show that we can easily distinguish a complexity-theoretic pseudorandom ensemble from the uniform ensemble. Both notions of pseudorandom strings are uniformly unpredictable; in contrast with pseudorandom strings, complexity-theoretic pseudorandom strings are not polynomial-time unpredictable.

**Key words:** pseudorandom generators, Kolmogorov complexity, unpredictability.

## 1 Introduction

There are two possible approaches to define the concept of randomness. The "ontological" approach looks at the "simplest description" of a string and declares random a string which has roughly the same length as its simplest description. Algorithmic information theory—initiated by Solomonoff [14], Kolmogorov [10], and Chaitin [4]—defines the simplest description of a string $x$ by the minimal input necessary to a universal algorithm to produce $x$. Depending upon the choice of the universal algorithm, two theories have emerged: Kolmogorov-Chaitin theory in which one uses a universal Turing machine and Chaitin theory relying on a self-delimiting universal Turing machine (see, e.g., [5]). Only the second theory is compatible with a theory of random infinite sequences. The first theory has been relativized (in time or space); it led to some complexity-theoretic definitions of pseudorandom strings. These notions have been very useful in many places (see [9] for a recent survey), but as Goldreich [8] observed, not in designing pseudorandom generators.

Cryptography suggests an alternative "behaviouristic" approach to pseudorandomness. Instead of considering the "explanation" of a phenomenon, it takes into account the phenomenon's effect on the environment. A string is said to be pseudorandom if no efficient observer can distinguish it from a uniformly chosen string of the same length. The underlying postulate is that objects that cannot be told apart by efficient procedures are considered equivalent. This approach naturally leads to the concept of pseudorandom generator, which is fundamental for cryptography.

Our aim is to contrast these two definitions of pseudorandom strings. For example, we show that we can easily distinguish a complexity-theoretic pseudorandom ensemble from the uniform ensemble. Both notions of pseudorandom strings are uniformly unpredictable; in contrast with pseudorandom strings, complexity-theoretic pseudorandom strings are not polynomial-time unpredictable.

We close this section by introducing some notation we will use. The set of nonnegative integers is denoted by $\mathcal{N}$. By $\{0,1\}^*$ we denote the set of (finite) binary strings; $\{0,1\}^n$ is the set of binary strings of length $n$. The length of a string $x$ is denoted by $|x|$. For a string $x \in \{0,1\}^*$ and an integer number $n \geq 1$, $x[1..n]$ denotes the initial segment of length $n$ of $x$ ($x[1..n] = x$ if $|x| \leq n$) and $x[i]$ denotes the $i$-th bit of $x$, i.e., $x[1..n] = x[1]\ldots x[n]$.

# 2 Computational indistinguishability

Computational indistinguishability is a fundamental concept in cryptography. The following paragraph is quoted from page 87 of [8]:

> The concept of efficient computation leads naturally to a new kind of equivalence between objects. Objects are considered to be computationally equivalent if they can not be told apart by any efficient procedure. Considering indistinguishable objects as equivalent is one of the basic paradigms of both science and real-life situations. Hence, we believe that the notion of computational indistinguishability is fundamental.

Two distributions are called computationally indistinguishable if no efficient algorithm can tell them apart. Given an efficient algorithm $D$, we consider the probability that $D$ accepts (e.g., outputs 1 on input) a string taken from the first distribution. Likewise, we consider the probability that $D$ accepts a string taken from the second distribution. If these two probabilities are close, we say that $D$ does not distinguish the two distributions.

Typically, an *ensemble* of the form $X = \{X_n\}_{n\in\mathcal{N}}$ has each $X_n$ ranging over strings of length $n$. We will use $U = \{U_n\}_{n\in\mathcal{N}}$ to denote the uniform ensemble, that is, $U_n$ denotes a random variable uniformly distributed over $\{0,1\}^n$.

**Definition 2.1** ([8]) *Two ensembles, $X = \{X_n\}_{n\in\mathcal{N}}$ and $Y = \{Y_n\}_{n\in\mathcal{N}}$, are* indistinguishable in polynomial-time *if for every probabilistic polynomial-time algorithm $D$, every polynomial $p(\cdot)$, and all sufficiently large $n$ such that the following two conditions are satisfied*

$$\sum_{x\in\{0,1\}^n} Prob(X_n = x) \neq 0 \quad\text{and}\quad \sum_{x\in\{0,1\}^n} Prob(Y_n = x) \neq 0,$$

*the following inequality holds:*

$$|Prob(D(X_n) = 1) - Prob(D(Y_n) = 1)| < \frac{1}{p(n)}.$$

The probabilities in the above definition are taken over the corresponding random variables $X_i$ (or $Y_i$) and the internal coin tosses of the algorithm $D$.

**Definition 2.2** *([8]) Let $U = \{U_n\}_{n\in\mathcal{N}}$ be the uniformly distributed ensemble, and $X = \{X_n\}_{n\in\mathcal{N}}$ be an ensemble. The ensemble $X$ is called* pseudorandom *if $X$ and $U$ are indistinguishable in polynomial-time.*

**Definition 2.3** *([8]) A* pseudorandom generator *is a deterministic polynomial-time algorithm $G$ from strings to strings satisfying the following two conditions:*

1. *There exists a function $l : \mathcal{N} \to \mathcal{N}$ such that $l(n) > n$ for all $n \in \mathcal{N}$, and $|G(x)| = l(|x|)$, for all $x \in \{0,1\}^*$.*
2. *The ensemble $\{G(U_n)\}_{n\in\mathcal{N}}$ is pseudorandom.*

# 3 No complexity-theoretic pseudorandom ensemble is pseudorandom

Let $\text{RAND}_c = \cup_{n\in\mathcal{N}}\text{RAND}_{c,n}$ and $\text{RAND}^t_c = \cup_{n\in\mathcal{N}}\text{RAND}^t_{c,n}$ be the sets of Kolmogorov $c$-random and Kolmogorov $t$-time bounded $c$-random strings respectively, where $c \geq 1$ and $t : \mathcal{N} \to \mathcal{N}$ is some time-constructible function such that $t(n) \geq n^2$ for all $n \in \mathcal{N}$. That is, for a universal Turing machine $M$, let

$$\text{RAND}_{c,n} = \{x \in \{0,1\}^n : \text{ if } M(y) = x \text{ then } |y| \geq |x| - c\}$$

and

$$\text{RAND}^t_{c,n} = \left\{ x \in \{0,1\}^n : \begin{array}{l} \text{if } M(y) = x \text{ and } M(y) \text{ halts in less} \\ \text{than } t(|x|) \text{ steps then } |y| \geq |x| - c \end{array} \right\}.$$

The strings in $\mathrm{RAND}_c$ (respectively $\mathrm{RAND}_c^t$) are called $c$-random (respectively $c$-pseudorandom). Let $R_c = \{R_{c,n}\}_{n\in\mathcal{N}}$ and $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ be two ensembles such that $R_{c,n}$ and $R_{c,n}^t$ are uniformly distributed over $\mathrm{RAND}_{c,n}$ and $\mathrm{RAND}_{c,n}^t$ respectively. Our first results show that these two ensembles are not pseudorandom.

**Theorem 3.1** *The ensemble* $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ *is not pseudorandom.*

*Proof.* Define a polynomial-time algorithm $D$ by letting

$$D(x) = \begin{cases} 0, & \text{if } x = 0^{\lfloor \log |x| \rfloor} y \text{ for some } y \in \{0,1\}^*, \\ 1, & \text{otherwise.} \end{cases}$$

It is straightforward to show that

$$\mathrm{RAND}_{c,n}^t \cap \{x \in \{0,1\}^* : x = 0^{\lfloor \log |x| \rfloor} y \text{ for some } y \in \{0,1\}^*\} = \emptyset$$

for sufficiently large $n$. Hence $Prob(D(R_{c,n}^t) = 1) = 1$ and $Prob(D(U_n) = 1) = 1 - 2^{\lfloor \log n \rfloor - n}$, for sufficiently large $n$. That is,

$$\Big| Prob(D(R_{c,n}^t) = 1) - Prob(D(U_n) = 1) \Big| = 2^{\lfloor \log n \rfloor - n} \geq \frac{1}{n}.$$

This shows that the ensembles $\{R_{c,n}^t\}_{n\in\mathcal{N}}$ and $\{U_n\}_{n\in\mathcal{N}}$ are distinguishable in polynomial-time, hence $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ is not pseudorandom. ∎

**Theorem 3.2** *The ensemble* $R_c = \{R_{c,n}\}_{n\in\mathcal{N}}$ *is not pseudorandom.*

*Proof.* The proof is similar to the proof of Theorem 3.1. ∎

# 4 Unpredictability

In this section we will show that $c$-random strings, $c$-pseudorandom strings, and pseudorandom strings are uniformly unpredictable. In contrast with pseudorandom strings, complexity-theoretic pseudorandom strings are not polynomial-time unpredictable.

## 4.1 Uniform unpredictability

One of the fundamental properties of random strings is the unpredictability of the $i$-th bit from the first $i-1$ bits of the sequence (see [15]). A weaker property, has been discussed in [3]: strings in $\mathrm{RAND}_c$ are normal.

**Definition 4.1** *Let $p(\cdot)$ be a given polynomial. An ensemble $X = \{X_n\}_{n\in\mathcal{N}}$ is called* uniformly unpredictable in polynomial-time *if for every polynomial-time algorithm $D : \{0,1\}^* \to \{0,1\}$, there is a constant $n_0$ such that for all $n \geq n_0$, a string $x \in X_n$ satisfies the following condition (1) with a probability of at least $1 - \frac{1}{p(n)}$:*

$$\left| \frac{\|\{i < n : D(x[1..i-1]) = x[i]\}\|}{n} - \frac{1}{2} \right| < \sqrt{\frac{\log n \log\log n}{n}}. \tag{1}$$

Note that, due to the law of the iterated logarithm, in (1) the bound $\sqrt{\frac{\log n \log\log n}{n}}$ cannot be strengthened to $\frac{1}{p(n)}$, for some polynomial $p(\cdot)$. In [15] it is shown that the law of the iterated logarithm holds for infinite pseudorandom sequences.

Now we show that both type pseudorandom ensembles are uniformly unpredictable in polynomial-time. We need for the proof Chernoff's Bound.

**Chernoff's Bound**. (See, e.g., Feller [6]) Let $X_1, X_2, \ldots, X_n$ be independent 0-1 random variables so that $Prob(X_i = 1) = \frac{1}{2}$, for each $i$. Then, for all $0 < \delta < \frac{1}{4}$, the following condition holds:

$$Prob\left( \left| \frac{\sum_{i=1}^n X_i}{n} - \frac{1}{2} \right| \geq \delta \right) < 2 \cdot e^{-2n\delta^2}. \tag{2}$$

**Corollary 4.2** *For each $n$ and $0 < \delta < \frac{1}{4}$, we have*

$$\left\| \left\{ x \in \{0,1\}^n : \left| \frac{\sum_{i=1}^n x[i]}{n} - \frac{1}{2} \right| \geq \delta \right\} \right\| < 2^{n+1} \cdot e^{-2n\delta^2}.$$

*Proof.* It follows from Chernoff's bound (2). ■

**Lemma 4.3** *Let $U = \{U_n\}_{n\in\mathcal{N}}$ be the uniform ensemble, $D : \{0,1\}^* \to \{0,1\}$ be a polynomial-time algorithm, and $\{A_n^D\}_{n\in\mathcal{N}}$ be a sequence of sets of strings defined as follows:*

$$A_n^D = \{x \in \{0,1\}^n : (1) \text{ does not hold for } x\}. \tag{3}$$

*Then $A^D = \cup_{n=1}^{\infty} A_n^D$ is a polynomial-time computable set and*

$$\|A_n^D\| \leq 2^{n+1-2\log e \log n \log\log n}$$

*for sufficiently large $n$.*

*Proof.* It is straightforward to check that $A^D$ is polynomial-time computable. Define a function $F$ from strings to strings by

$$F(x) = (D(\lambda) \oplus x[1])(D(x[1..1]) \oplus x[2]) \ldots (D(x[1..n-1]) \oplus x[n])$$

for each $x \in \{0,1\}^n$, where $\lambda$ is the empty string. Let $B_n = \left\{F(x) : x \in A_n^D\right\}$. Then it is straightforward that for each $x \in B_n$, we have

$$\left|\frac{\sum_{i=1}^n x[i]}{n} - \frac{1}{2}\right| \geq \sqrt{\frac{\log n \log\log n}{n}}.$$

Now let $\delta = \sqrt{\frac{\log n \log\log n}{n}}$. Then, by Corollary 4.2, we derive the following bound for the cardinality of $A_n^D$:

$$\begin{aligned} \|A_n^D\| = \|B_n\| &\leq 2^{n+1} \cdot e^{-\frac{2n \log n \log\log n}{n}} \\ &= 2^{n+1} \cdot e^{-2\log n \log\log n} \\ &= 2^{n+1-2\log e \log n \log\log n}. \end{aligned}$$

■

**Theorem 4.4** *The ensemble* $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ *is uniformly unpredictable in polynomial-time, where* $t(n) \geq 2^{2n}$ *is some time-constructible function.*

*Proof.* Let $D : \{0,1\}^* \to \{0,1\}$ be a polynomial-time algorithm, and $\{A_n^D\}_{n\in\mathcal{N}}$ as in Lemma 4.3. Since any member $x$ of the set $A_n^D$ can be calculated uniquely in time $2^{2n}$ if we are given the polynomial-time algorithm $D$ and the position of $x$ in $A_n^D$ expressed as an $n-[2\log e \log n \log\log n]$ bit string. It follows that $A_n^D \cap \mathrm{RAND}_{c,n}^t = \emptyset$, for sufficiently large $n$. This means that the ensemble $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ is uniformly unpredictable in polynomial-time. ■

**Theorem 4.5** *The ensemble* $R_c = \{R_{c,n}\}_{n\in\mathcal{N}}$ *is uniformly unpredictable in polynomial-time.*

*Proof.* The proof is similar to the proof of Theorem 4.4. ■

**Theorem 4.6** *Every pseudorandom ensemble* $X = \{X_n\}_{n\in\mathcal{N}}$ *is uniformly unpredictable in polynomial-time.*

*Proof.* For the sake of contradiction, we assume that $X$ is not uniformly unpredictable in polynomial-time. That is, there is a polynomial-time algorithm $D : \{0,1\}^* \to \{0,1\}$ and a polynomial $p_0(\cdot)$ such that the following condition holds for infinitely many $n$:

$$\sum_{x \in A_n^D} Prob(X_n = x) > \frac{1}{p_0(n)}, \tag{4}$$

where $\{A_n^D\}_{n\in\mathcal{N}}$ is defined in Lemma 4.3. Now we define a polynomial-time computable function $D'$ by letting

$$D'(x) = \begin{cases} 1, & x \in A_n^D \text{ for some } n \in \mathcal{N}, \\ 0, & \text{otherwise.} \end{cases}$$

By virtue of the definition of $D'$, we have the following equality:

$$\begin{aligned} & Prob(D'(X_n) = 1) - Prob(D'(U_n) = 1) \\ = \quad & \textstyle\sum_{x\in A_n^D} Prob(X_n = x) - \sum_{x\in A_n^D} Prob(U_n = x). \end{aligned}$$

Hence, by Lemma 4.3 and (4), the following inequality holds for sufficiently large $n$:

$$\begin{aligned} |Prob(D'(X_n) = 1) - Prob(D'(U_n) = 1)| \quad &\geq \quad \frac{1}{p_0(n)} - e^{-2\log n \log\log n} \\ &= \quad \frac{1}{p_0(n)} - \frac{1}{n^{2\log e \log\log n}} \\ &\geq \quad \frac{1}{2p_0(n)}. \end{aligned}$$

This contradicts with the fact that $X$ and $U$ are indistinguishable in polynomial-time. ∎

**Corollary 4.7** *The uniform ensemble $U = \{U_n\}_{n\in\mathcal{N}}$ is uniformly unpredictable in polynomial-time.*

*Proof.* This follows from Theorem 4.6. ∎

Since the ensemble $R_c = \{R_{c,n}\}_{n\in\mathcal{N}}$ is uniformly unpredictable in polynomial-time (cf. Theorem 4.5) but not pseudorandom (cf. Theorem 3.2), the converse of Theorem 4.6 is not true.

**Corollary 4.8** *Let $G$ be a pseudorandom generator. Then the ensemble $\{G(U_n)\}_{n\in\mathcal{N}}$ is uniformly unpredictable in polynomial-time.*

*Proof.* This follows from Theorem 4.6. ■

Theorem 4.8 shows that given a pseudorandom generator $G$, and a truly random input $x$, the output $G(x)$ is unpredictable in polynomial-time with high probability, though $G(x)$ is not $c$-pseudorandom.

## 4.2 Cryptographic unpredictability

**Definition 4.9** ([16]) *An ensemble $X = \{X_n\}_{n\in\mathcal{N}}$ is called* unpredictable in polynomial-time *if for every probabilistic polynomial-time algorithm $D$, every polynomial $p(\cdot)$, and all sufficiently large $n$, the following condition is satisfied.*

$$Prob\,(D(X_n) = next_D(X_n)) \leq \frac{1}{2} + \frac{1}{p(n)},$$

*where $next_D(x)$ returns the $(i+1)$-th bit of $x$ if $D$ on input $x$ reads only $i < |x|$ bits of $x$, and returns a uniformly chosen bit otherwise (i.e., in case $D$ read the entire string $x$).*

**Theorem 4.10** ([16], [2]) *An ensemble $X = \{X_n\}_{n\in\mathcal{N}}$ is pseudorandom if and only if it is unpredictable in polynomial-time.*

**Corollary 4.11** *Neither the ensemble $R_c = \{R_{c,n}\}_{n\in\mathcal{N}}$ nor the ensemble $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ is unpredictable in polynomial-time.*

*Proof.* This follows from Theorems 3.1, 3.2, and 4.10. ■

# 5 Strong unpredictability

In the view of Definition 4.1, an ensemble $X = \{X_n\}_{n\in\mathcal{N}}$ is uniformly unpredictable in polynomial-time if for every polynomial-time algorithm $D : \{0,1\}^* \to \{0,1\}$ and sufficiently large $n$, a string $x \in X_n$ satisfies (1) with a probability of at least $1-\frac{1}{p(n)}$. If we replace the probability $1-\frac{1}{p(n)}$ with 1, then we obtain a stronger definition.

**Definition 5.1** *An ensemble $X = \{X_n\}_{n\in\mathcal{N}}$ is called* strongly unpredictable in polynomial-time *if for every polynomial-time algorithm $D : \{0,1\}^* \to \{0,1\}$, there is a constant $n_0$ such that for all $n \geq n_0$ and all strings $x$ such that $Prob(X_n = x) > 0$, condition (1) holds.*

The proof of Theorem 4.4 shows that the ensemble $R_c^t = \{R_{c,n}^t\}_{n\in\mathcal{N}}$ is strongly unpredictable in polynomial-time. However, pseudorandom ensembles are not necessarily strongly unpredictable in polynomial-time. For example, the uniform ensemble $U = \{U_n\}_{n\in\mathcal{N}}$ is not strongly unpredictable in polynomial-time. As another example, we show that the ensemble $\{G(U_n)\}_{n\in\mathcal{N}}$ is not strongly unpredictable in polynomial-time where $G$ is the BBS [1] pseudorandom generator.

**Example 1** *Let both $p$ and $q$ be distinct primes congruent to 3 mod 4, $N = pq$, and $l(n) > n$ be a polynomial. For each number $x < N$ and $i \leq l(\log N)$, let $x_{-1} = x$, $x_{i+1} = x_i^2 \bmod N$ and $b_i$ =parity($x_i$) where parity($y$) denotes the least significant bit of $y$. Then the BBS [1] pseudorandom generator is defined as $G(x) = b_0 \dots b_{l(\log N)}$. It is clear that $G(0) = 0\dots 0$. Whence $\{G(U_n)\}_{n\in\mathcal{N}}$ is not strongly unpredictable in polynomial-time.*

After the above discussion, one may wonder whether there exists an ensemble which is both pseudorandom random and strongly unpredictable in polynomial-time. The following theorem gives an affirmative answer.

**Theorem 5.2** *Let $D_1, D_2, \dots$ be a uniform enumeration (that is, $D_i(x)$ is computable in time $2^{|x|+i}$) of all polynomial-time algorithms, and $A_n^{D_i}$ be defined in Lemma 4.3. Then the ensemble $X = \{X_n\}_{n\in\mathcal{N}}$ is both pseudorandom and strongly unpredictable in polynomial-time, where $X_n$ is a random variable uniformly distributed over $\{0,1\}^n \setminus (\cup_{i=1}^{\lfloor\log\log n\rfloor} A_n^{D_i})$.*

*Proof.* It is left as an excise. ■

However, the following question remains open.

**Question 1**. For a pseudorandom generator $G$, is the ensemble $\{G(R_{c,n})\}_{n\in\mathcal{N}}$ strongly unpredictable in polynomial-time?

If the answer to the above question is positive, then we get a characterization of pseudorandom generators. That is, for a pseudorandom generator $G$ and a truly random input $x \in \mathrm{RAND}_c$, the output $G(x)$ satisfies the condition (1). This coincides with our intuition that the $i$-th bit of a pseudorandom string should not be predictable from its first $i-1$ bits. However, the answer to Question 1 may be negative; in this case we suggest the following alternative definitions for pseudorandom generators.

**Definition 5.3 (Suggested new definition 1).** *A* pseudorandom generator *is a deterministic polynomial-time algorithm $G$ satisfying the following three conditions:*

1. *There exists a function* $l : \mathcal{N} \to \mathcal{N}$ *so that* $l(n) > n$ *for all* $n \in \mathcal{N}$, *and* $|G(s)| = l(|s|)$ *for all* $s \in \{0,1\}^*$.
2. *The ensemble* $\{G(U_n)\}_{n\in\mathcal{N}}$ *is pseudorandom.*
3. *The ensemble* $\{G(R_{c,n})\}_{n\in\mathcal{N}}$ *is strongly unpredictable in polynomial-time.*

**Definition 5.4 (Suggested new definition 2).** *A* pseudorandom generator *is a deterministic polynomial-time algorithm* $G$ *satisfying the following two conditions:*

1. *There exists a function* $l : \mathcal{N} \to \mathcal{N}$ *so that* $l(n) > n$ *for all* $n \in \mathcal{N}$, *and* $|G(s)| = l(|s|)$ *for all* $s \in \{0,1\}^*$.
2. *The ensemble* $\{G(R_{c,n})\}_{n\in\mathcal{N}}$ *is strongly unpredictable in polynomial-time.*

As a summary, we list in Table 1 a comparison of the complexity-theoretic pseudorandom ensembles and cryptographic pseodorandom ensembles.

Table 1: A comparison of two approaches to pseudorandomness

| | complexity approach ($R_c$ and $R_c^t$) | cryptographic approach ($U, \{G(U_n)\}_{n\in\mathcal{N}}$, etc.) |
|---|---|---|
| indistinguishable from $\{U_n\}_{n\in\mathcal{N}}$? | no | yes |
| uniformly unpredictable? | yes | yes |
| cryptographicly unpredictable? | no | yes |
| strongly unpredictable? | yes | yes **or** no |

# 6 Pseudorandomness in practice

In the previous section, we recommended new definitions of pseudorandmness in terms of strong unpredictability and Kolmogorov complexity. In practice, we may not need such stronger definition. For example, instead of considering the ensemble $R_c^t$, it is practically sufficient to consider the sequences which withstand the following five basic tests (see, e.g., [13]).

1. Frequency test (mono-bit test). The purpose of this test is to determine whether the numbers of 0's and 1's in a sequence are approximately the same, as would be expected for a random sequence.

2. Serial test (two-bit test). The purpose of this test is to determine whether the numbers of occurrences of each two-bit sequence are approximately the same.

3. Poker test. This is a generalization of the frequency test (see [13] for details).

4. Runs test. The purpose of the runs test is to determine whether the numbers of runs (of either zeros or ones) of various lengths in the sequence is as expected for a random sequence.

5. Autocorrelation test. The purpose of this test is to check for correlations between the sequence and (non-cyclic) shifted versions of it.

Following the suggestion in FIPS 140-1 (see [7]), we may call a specific sequence $x$ *FIPS-pseudorandom* if it withstands the following four tests: monobit test, poker test, runs test, and long run test (see [7] for details). Let $FIPS_n$ be the set of all $n$-length FIPS-pseudorandom sequences. Then in practice we may consider the following definition for pseudorandom generators.

**Definition 6.1** *A* pseudorandom generator *is a deterministic polynomial-time algorithm G satisfying the following three conditions:*

1. *There exists a function* $l : \mathcal{N} \to \mathcal{N}$ *so that* $l(n) > n$ *for all* $n \in \mathcal{N}$*, and* $|G(s)| = l(|s|)$ *for all* $s \in \{0,1\}^*$.

2. *The ensemble* $\{G(U_n)\}_{n\in\mathcal{N}}$ *is pseudorandom.*

3. *The ensemble* $\{G(FIPS_n)\}_{n\in\mathcal{N}}$ *is strongly unpredictable in polynomial-time.*

Alternatively, we may also consider the following weaker definition.

**Definition 6.2** *A* pseudorandom generator *is a deterministic polynomial-time algorithm G satisfying the following conditions:*

1. *There exists a function* $l : \mathcal{N} \to \mathcal{N}$ *so that* $l(n) > n$ *for all* $n \in \mathcal{N}$*, and* $G(FIPS_n) \subseteq FIPS_{l(n)}$ *for all* $n \in \mathcal{N}$.

2. *The ensemble* $\{G(U_n)\}_{n\in\mathcal{N}}$ *is pseudorandom.*

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